

McOE: A Family of Almost Foolproof On-Line Authenticated Encryption Schemes

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Abstract. *On-Line Authenticated Encryption* (OAE) combines privacy with data integrity and is on-line computable. Most block-cipher-based schemes for Authenticated Encryption can be run on-line and are provably secure against *nonce-respecting* adversaries. But they fail badly for more general adversaries. This is not a theoretical observation only – in practice, the reuse of nonces is a frequent issue¹.

In recent years, cryptographers developed *misuse-resistant* schemes for Authenticated Encryption. These guarantee excellent security even against general adversaries which are allowed to reuse nonces. But they can not perform on-line encryption.

This work introduces a new family of OAE schemes –called McOE– dealing both with nonce-respecting *and* with general adversaries. Furthermore, we present two block-cipher-based family members, *i.e.*, McOE-X and McOE-G. In contrast to other published OAE, they provably guarantee reasonable security against general adversaries as well as standard security against nonce-respecting adversaries.

Keywords: authenticated encryption, on-line encryption, provable security, misuse-resistant

¹ A prominent example is the PlayStation 3 'jailbreak' [20], where application developers used a constant that was actually supposed to be a nonce for a digital signature scheme.

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Construction

1 Introduction

On-Line Authenticated Encryption (OAE). Application software often requires a network channel that guarantees both privacy and authenticity of the data being communicated between two parties. Cryptographic schemes able to meet twofold these goals are commonly referred to as Authenticated Encryption (AE) schemes.

The ISO/IEC 19772:2009 standard for AE [21] defines generic composition (Encrypt-then-MAC [3]) and five dedicated AE schemes: OCB2 [38], SIV [41] (denoted as “Key Wrap” in [21]), CCM [14], EAX [7], and GCM [34]. To integrate an AE-secure channel most seamlessly into a typical software architecture, application developers expect it to encrypt in an *on-line* manner, *i.e.*, the i -th ciphertext block can be written before the $(i+1)$ -th plaintext block is read. Whereas off-line encryption, where either the entire plaintext must be known in advance or read more than once, is an encumbrance to software architects.

Nonces and their reuse. Goldwasser and Micali [18] formalized encryption schemes as stateful or probabilistic, because otherwise important security properties are lost. Rogaway [37,39,40] proposed a unified point of view by always defining a cryptographic scheme as a deterministic algorithm that takes user-supplied nonce (a number used once). Then the application programmer – and not the encryption scheme – is responsible for flipping coins or maintaining state. This reflects cryptographic practice since the algorithm itself is often implemented by a multi-purpose cryptographic library which is more or less application-agnostic.

In theory, the concept of nonces is simple. In practice, it is challenging to ensure that a nonce is *never* reused. Flawed implementations of nonces are ubiquitous [10,20,28,44,45]. Apart from implementation failures, there are fundamental reasons why software developers cannot always prevent nonce-reuse. A persistently stored counter, which is increased and written back each time a new nonce is needed, may be reset by a backup – usually after a previous data loss. Similarly, the internal and persistent state of an application may be duplicated when a virtual machine is cloned, etc.

Related Work and Our Contribution. We aim to achieve *simultaneously*: security against nonce-reusing adversaries (sometimes also called nonce-misusing adversaries) *and* support for on-line-encryption in terms of an AE scheme. Apart from generic composition (Encrypt-then-Mac, EtM), none of the ISO/IEC 19772:2009 schemes – in fact, no previously published AE scheme at all – achieves both goals, (cf. Table 1). In this table, we classify a variety of provably secure block-cipher-based AE schemes with respect to their on-line-ability and against which adversaries (nonce-respecting versus -reusing) they are proven secure.

Since EtM is not a concrete scheme but a generic construction technique, there are some challenges left in order to make it fully on-line secure: First, an appropriate on-line cipher has to be chosen. Second, a suitable, on-line-computable, secure, and deterministic MAC must be selected. And, third, the EtM scheme requires at least two *independent* keys to be secure. Since two schemes are used in parallel, it is likely to squander resources in terms of run time and – important for hardware designers – in terms of space. Since EtM first has to be turned into an OAE scheme by making the appropriate choices, we do not include it in our analysis.

As it turned out, we actually found nonce-reuse attacks for *all* of those schemes, cf. Table 2 and Appendix A.

In this paper we present a new family of on-line authenticated encryption schemes called MCOE. The general structure is based on the Tweak Chain Hash (TCH) from [31] which is adapted from the Matyas-Meyer-Oseas (MMO) construction [?]. We introduce two members of the MCOE family – called MCOE-X and MCOE-G. That are both are able to fill the gap in

secure ...	against nonce-respecting adversaries	ag. nonce-reusing adversaries
on-line	CCFB[33] CHM[22] CIP[23] CWC[29] EAX[7] GCM[34] IACBC[26] IAPM[26] McOE-G McOE-X OCB1-3 [40,38,30] RPC[11] TAE[31] XCBC[17]	McOE-X (this paper) McOE-G (this paper)
off-line	BTM[24] CCM[14] HBS[25] SIV[41] SSH-CTR[36]	BTM[24] HBS[25] SIV[41]

Table 1. Classification of provably secure block-cipher-based AE Schemes. CCM and SSH-CTR are considered off-line because encryption requires prior knowledge of the message length. Note that the family of McOE schemes, because of being on-line, satisfies a slightly weaker security definition against nonce-reusing adversaries than SIV, HBS, and BTM.

	privacy attack workload	authenticity attack workload		privacy attack workload	authenticity attack workload
CCFB [33]	$O(1)$	$O(1)$	IAPM [26]	$O(1)$	$O(1)$
CCM [14]	$O(1)$	$\ll 2^{(n/2)}$ [15]	OCB1 [40]	$O(1)$	$O(1)$
CHM [22]	$O(1)$	$O(1)$	OCB2 [38]	$O(1)$	$O(1)$
CIP [23]	$O(1)$	$O(1)$	OCB3 [30]	$O(1)$	$O(1)$
CWC [29]	$O(1)$	$O(1)$	RPC [11]	$O(1)$	$O(1)$
EAX [7]	$O(1)$	$O(1)$	TAE [?]	$O(1)$	$O(1)$
GCM [34]	$O(1)$	$O(1)$	XCBC [17]	$O(2^{n/4})$?
IACBC [26]	$O(1)$	$O(1)$			

Table 2. Overview of our **nonce-reuse** attacks on published AE schemes, excluding SIV, HBS and BTM, which have been explicitly designed to resist nonce-reuse. Almost all attacks achieve an advantage close to 1. The “attack workload” covers the computational effort, amount of needed memory as well as the time complexity. Details are given in Appendix A.

the upper-right position of Table 1. We argue that closing this gap is both practically relevant and theoretically interesting.

Initial Value (IV)-based AE schemes which provide security against repeated IV’s have been addressed by Rogaway and Shrimpton in [41]. Furthermore, the authors shaped the notion of *misuse-resistance* and they proposed SIV as a solution. SIV and related schemes (HBS [25] and BTM [24]) actually provide excellent security against nonce-reusing adversaries, though, there are other potential cases of misuse (cf. Appendix A.3). Their major disadvantage is that they are inherently off-line: For encryption, one must either keep the entire plaintext in memory, or read the plaintext twice.

Ideally, an adversary seeing the encryptions of two (equal-length) plaintexts P_1 and P_2 cannot even decide if $P_1 = P_2$ or not. When using a nonce more than once, deciding about $P_1 = P_2$ is easy. SIV and its relatives ensure that nothing else is feasible for nonce-reusing adversaries. In the case of on-line encryption, where the first few bits of the encryption of a lengthy message must not depend on the last few bits of that message, there unavoidably is something beyond $P_1 = P_2$. The adversary can compare any two ciphertexts for their longest common prefix, and then conclude about common prefixes of the secret plaintexts. Our notion of *misuse-resistance* means that this is all the adversary can gain. Even in the case of a nonce-reuse:

1. The adversary cannot do anything beyond determining the length of common plaintext prefixes and
2. the scheme still provides the usual level of authenticity for AE (INT-CTXT).

The first property is common for on-line ciphers/permutations (OPRP) [1]. Recently, [43] studied the design of on-line ciphers from tweakable block ciphers, bearing some similarities to our

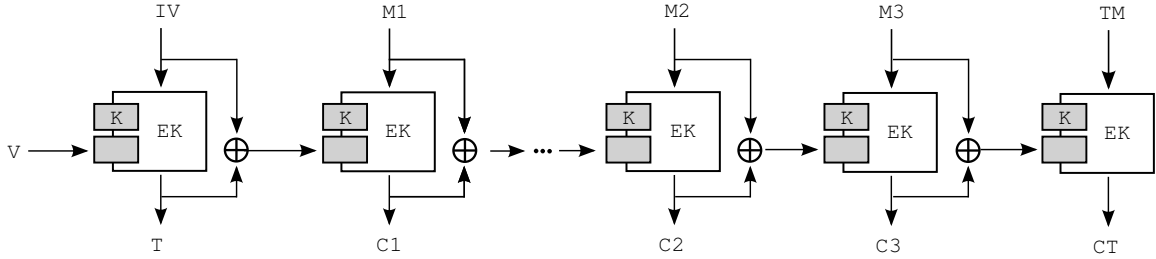


Fig. 1. The generic MCOE construction, where \tilde{E} denotes a tweakable block cipher.

approach, especially to TC3. In contrast to the MCOE family, the constructions from [43] provide no authentication.

Design Principles for AE Schemes. The question of how to provide authenticated encryption (without stating that name) when given a secure on-line cipher is studied in [2], the revised and full version of [1]. The first approach in [2] only provides security if all messages are of the same length. The second approach repairs that by prepending the message’s length to the message, at the cost of being off-line since the message length must be known at the beginning of the encryption process. Their approach is to prepend and append a random W to a message M and then to perform the on-line encryption of $(W||M||W)$. This looks promising, but the same W is used for two different purposes, putting different constraints on the generation of W . For privacy, it suffices that W behaves like a nonce, not requiring secrecy or unpredictability. Even if W is not a nonce, but the same W is used for the encryption of several messages, all the adversary can determine are the lengths of common plaintexts prefixes, as we required for nonce-reuse. On the other hand, authenticity actually assumes a *secret or unpredictable* W rather than a nonce. If the adversary can guess W before choosing a message, she asks for the authenticated encryption of $(M||W)$. Then, she can predict the authenticated encryption of M without actually asking for it.

The MCOE family replaces the “random” W by a proper nonce and the *key-dependent* tag computation value τ , performing a nonce-dependent on-line encryption of $(M||\tau)$. The encryption can also depend on some associated data, which turns MCOE into a family of schemes for OAEAD (*On-Line Authenticated Encryption with Associated Data*).

Roadmap. In Section 2 we describe the basic design principle of MCOE. Then, we present the members of the MCOE family. Furthermore, we introduce its concrete instances, and provide performance data when instantiated with either AES-128 or Threefish-512 as underlying block cipher. Section 3 deals with general notions and definitions, and Section 4 defines the security of OAE. The main result of the paper, the security proof of the generic MCOE scheme and its analysis is presented in Section 5. In Sections 6 and 7 we show the security of MCOE-X and MCOE-G, respectively.

The discussion in Section 8 concludes the paper. The appendix deals with misuse-attacks against published AE schemes, and provides proof supplements.

2 Practical On-Line Authenticated Encryption using AES and Threefish

2.1 Generic Construction of the MCOE Family (without Tag-splitting)

Our design goal was to build a misuse-resistant on-line authenticated-encryption scheme, which follows the on-line-permutation approach discussed by Bellare *et al.* in [1]. Therefore, our generic

MCOE structure (see Figure 1) is based on TC3, which is an on-line encryption scheme presented by Rogaway and Zhan in [43]. Like TC3, our scheme is based on a tweakable block cipher – called \tilde{E}_K – but is stateful regarding to the usage of a nonce. Additionally, we expanded it to a full-fledged authenticated encryption scheme with an additional effort of only two tweakable block cipher calls.

We also introduce the *tag-splitting* (TS) method for processing messages whose length is not a multiple of the block length. Without TS, we would have to pad such messages and then encrypt the padded messages – resulting in an expanded ciphertext. TS is similar to a well-known length-preserving method called *ciphertext stealing* (CTS), *e.g.*, [13]. Note, that CTS requires to process the final block before its predecessors, which is not possible for MCOE.

The encryption and decryption of the generic construction of MCOE without TS can be described by the following two algorithms.

Definition 1 (Generic McOE Scheme without Tag-Splitting). *Let $\Pi = \{\mathcal{K}, \mathcal{E}, \mathcal{D}\}$ be an authenticated-encryption scheme, where \mathcal{K} denotes the key-derivation function, \mathcal{E} the encryption function, and \mathcal{D} the decryption function. Let $\tilde{E} \in \text{Block}(n, n, n)$ be a tweakable block cipher. Furthermore, let H be the header with $H \in D_n^{L_H}$, M be the message with $M \in D_n^L$ for some integer L , T be the authentication tag with $T \in D_n$, and C be the ciphertext with $C \in D_n^L$. Then \mathcal{E} and \mathcal{D} of the MCOE family are given by the algorithms **EncryptAuthenticate** and **DecryptAuthenticate**, respectively. Here, the encryption function takes a header H and a message M returning a ciphertext C and a tag T . The decryption function takes a header H , a ciphertext C , and a tag T , and returns either a plaintext M or the fail symbol \perp .*

<p>EncryptAuthenticate(H, M)</p> <ol style="list-style-type: none"> 1. $U \leftarrow 0^n$ 2. for $i = 1, \dots, L_H - 1$ do $U \leftarrow \tilde{E}_K(U, H_i) \oplus H_i$ 3. $\tau \leftarrow \tilde{E}_K(U, H_{L_H})$ 4. $U \leftarrow \tau \oplus H_{L_H}$ 5. for $i = 1, \dots, L$ loop $C_i \leftarrow \tilde{E}_K(U, M_i)$ $U \leftarrow M_i \oplus C_i$ 6. $T \leftarrow \tilde{E}_K(U, \tau)$ 7. return (C_1, \dots, C_L, T) 	<p>DecryptAuthenticate(H, C, T)</p> <ol style="list-style-type: none"> 1. $U \leftarrow 0^n$ 2. for $i = 1, \dots, L_H - 1$ do $U \leftarrow \tilde{E}_K(U, H_i) \oplus H_i$ 3. $\tau \leftarrow \tilde{E}_K(U, H_{L_H})$ 4. $U \leftarrow \tau \oplus H_{L_H}$ 5. for $i = 1, \dots, L$ loop $M_i \leftarrow \tilde{E}_K^{-1}(U, C_i)$ $U \leftarrow M_i \oplus C_i$ 6. if $T = \tilde{E}_K(U, \tau)$ then return (M_1, \dots, M_L) else return \perp
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In the case when the header or the message length is not a multiple of the block size n , we recommend to use the secure 10*-padding. Furthermore, the header has to consist of at least one block, since the tag computation value τ depends on it. Hence, the whole header can be seen as a nonce. To fulfill the requirement of length-preserving encryption, we introduce the generic MCOE scheme using the TS method in the next section.

2.2 Generic Construction of the McOE Family (Tag-Splitting)

In Figure 2 you can see the generic MCOE scheme when using tag-splitting to be length preserving. Both the encryption and decryption process can be seen in the algorithms *EncryptAuthenticateSplitTag* (abbr. **EAST**) and *DecryptAuthenticateSplitTag* (abbr. **DAST**). Note that the additional effort to generate the tag – in comparison to MCOE without TS – is given by only one block cipher invocation (cf. line 7 and 9 in **EAST** and **DAST**, respectively).

Definition 2 (Generic McOE Scheme with Tag-Splitting). Let $\Pi = \{\mathcal{K}, \mathcal{E}, \mathcal{D}\}$ be an authenticated-encryption scheme, where \mathcal{K} denotes the key-derivation function, \mathcal{E} the encryption function, and \mathcal{D} the decryption function. Let $\tilde{E} \in \text{Block}(n, n, n)$ be a tweakable block cipher. Furthermore, let H be the header with $H \in D_n^{L_H}$, M be the message with $M \in D_n^L \parallel \{0, 1\}^{l^*}$ for some integers L and l^* , $0 < l^* < n$, T be the authentication tag with $T \in D_n$, and C be the ciphertext with $C \in D_n^L \parallel \{0, 1\}^{l^*}$. Then, the tag-splitting variants of \mathcal{E} and \mathcal{D} are given by the algorithms **EAST** and **DAST**, respectively. Here, the encryption function takes a header H and a message M , returning a ciphertext C and a tag T . The decryption function takes a header H , a ciphertext C , and a tag T and returns either a plaintext M or the fail symbol \perp .

EAST(H, M)

1. $U \leftarrow 0^n$
2. **for** $i = 1, \dots, L_H - 1$ **do**
 $U \leftarrow \tilde{E}_K(U, H_i) \oplus H_i$
3. $\tau \leftarrow \tilde{E}_K(U, H_{L_H})$
4. $U \leftarrow \tau \oplus H_{L_H}$
5. **for** $i = 1, \dots, L - 1$ **loop**
 $C_i \leftarrow \tilde{E}_K(U, M_i)$
 $U \leftarrow M_i \oplus C_i$
6. $M^* \leftarrow (M_L \parallel \tau[0 \dots n - l^* - 1])$
7. $M^* \leftarrow M^* \oplus \tilde{E}_K(1^n, |M_L|)$
8. $C^* \leftarrow \tilde{E}_K(U, M^*)$
9. Parse $C_L \parallel T[0 \dots n - l^* - 1] \leftarrow C^*$
10. $U \leftarrow M^* \oplus C^*$
11. $C^{**} \leftarrow \tilde{E}_K(U, \tau)$
12. $T[n - l^* \dots n - 1] \leftarrow C^{**}[0 \dots l^* - 1]$
13. **return** $(C_1, \dots, C_{L-1}, C_L^*, T)$

DAST(H, C, T)

1. $U \leftarrow 0^n$
2. **for** $i = 1, \dots, L_H - 1$ **do**
 $U \leftarrow \tilde{E}_K(U, H_i) \oplus H_i$
3. $\tau \leftarrow \tilde{E}_K(U, H_{L_H})$
4. $U \leftarrow \tau \oplus H_{L_H}$
5. **for** $i = 1, \dots, L - 1$ **loop**
 $M_i \leftarrow \tilde{E}_K^{-1}(U, C_i)$
 $U \leftarrow M_i \oplus C_i$
6. $C^* \leftarrow C_L \parallel T[0 \dots n - l^* - 1]$
7. $M^* \leftarrow \tilde{E}_K^{-1}(U, C^*)$
8. $U \leftarrow M^* \oplus C^*$
9. $M^* \leftarrow M^* \oplus \tilde{E}_K(1^n, |C_L|)$
10. Parse $M_L \parallel \tau'[0 \dots n - l^* - 1] \leftarrow M^*$
11. $T' \leftarrow \tilde{E}_K(U, \tau)$
12. **if** $\tau'[0 \dots n - l^* - 1] = \tau[0 \dots n - l^* - 1]$
and $T'[0 \dots l^* - 1] = T[n - l^* \dots n - 1]$
then return (M_1, \dots, M_L) **else return** \perp

Both schemes, with and without tag-splitting, are secure in the common CCA-setting, assuming a nonce-respecting adversary. In addition, they guarantee a certain amount of security in the nonce-misuse scenario, *i.e.*, indistinguishability from an on-line permutation and security against existential forgery attacks.

2.3 McOE-X

The first instance presented in this paper is called McOE-X, where the 'X' indicates the way of handling the tweak. This scheme uses an ordinary block cipher which is converted to a tweakable block cipher by XORing the tweak (*i.e.*, the chaining value) to the key K . A depiction of this instance is given in Figure 3.

The none-tag-splitting and tag-splitting modes of McOE-X can be described by the algorithms introduced in Section 2.1 (see Definition 1 and 2), where the tweakable block cipher \tilde{E}_K is defined by

$$\tilde{E}_K(U, M) := E_{K \oplus U}(M),$$

and E_K is a common block cipher, *e.g.*, AES, Serpent, or MARS. For performance testing we have implemented McOE-X using the block ciphers AES-128 and Threefish-512, resulting in the two practical instances McOE-X-AES and McOE-X-Threefish. Both implementations are

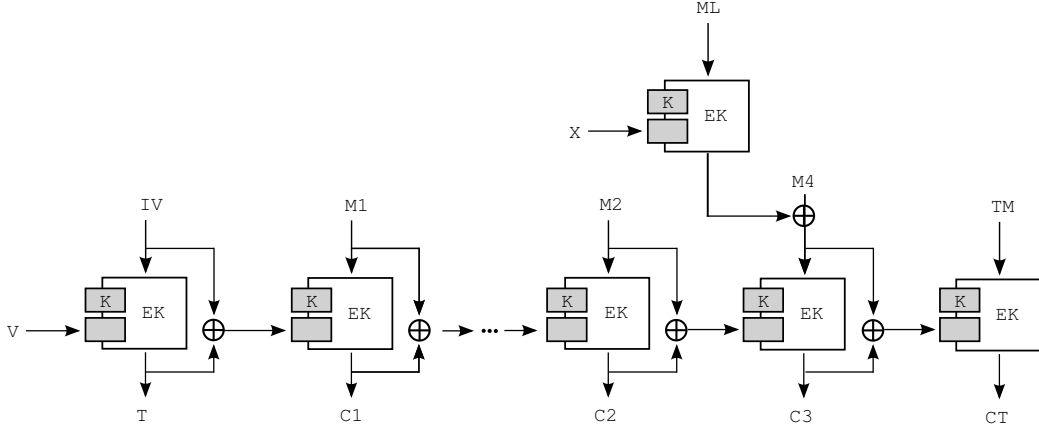


Fig. 2. The generic MCOE construction, where \tilde{E} denotes a tweakable block cipher. For simple reference, we denote T^α as the $(n - |C_L|)$ -bit string $T[0 \dots n - |C_L| - 1]$ and T^β as the $|C_L|$ -bit string $T[n - |C_L|, \dots, n]$. Additionally, we denote the corresponding strings $\tau^\alpha = \tau[0, \dots, |C_L| - 1]$ and $\tau^\beta = \tau[|C_L|, \dots, n - 1]$, respectively.

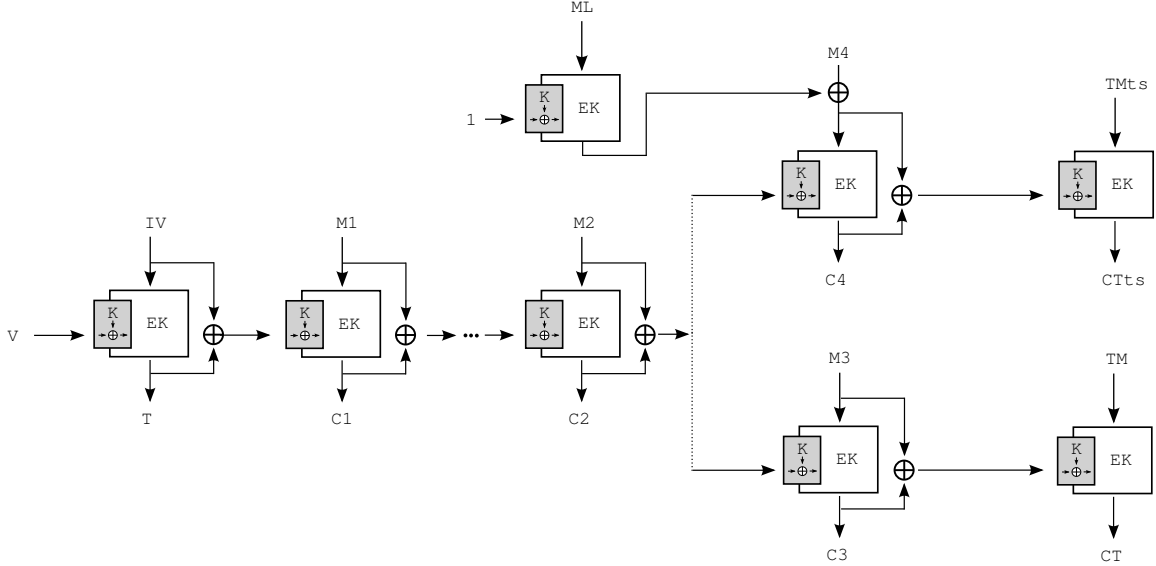


Fig. 3. The MCOE-X encryption process. In case that the message length is not a multiple of the block size, MCOE-X performs tag-splitting (upper variant). Otherwise, the tag can be computed without splitting (lower variant). The key used for the block cipher E is computed by the injective function $K \oplus U$ which is given the secret key K and the chaining value input U . The returned tag is the n -bit value T . The $(n - l)$ -bit value Z is discarded. The decryption process works in a similar way from 'left to right', only the block cipher component E is replaced by its counterpart E^{-1} apart from one exception: the first call computes τ .

easily extended to smoothly handle associated data, *i.e.*, data that is not encrypted but only authenticated. The security proofs considering associated data are given in Section 6.

The choice of Threefish-512 is based on two facts. First, it contains an agile key scheduler since it is optimized for hashing messages in the MMO (Matyas-Meyer-Oseas) mode. Second, it processes message blocks of 512 bit size, which results in less invocations of the block cipher E_K .

Remark. For MCOE-X we need related-key-resistance for the block cipher E since the adversary can 'partially control' some relations among keys used in the computation. We need this requirement only for MCOE-X and not for the two instances introduced in the next sections.

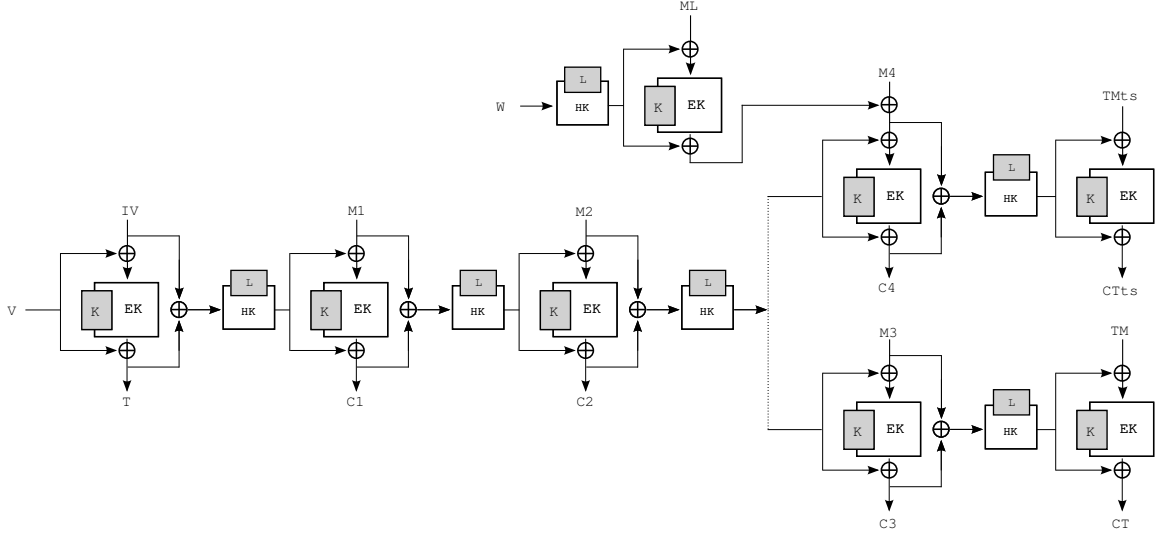


Fig. 4. The MCOE-G encryption process. If the message length is not a multiple of the block size, MCOE-G performs tag-splitting (upper variant), where T^α denotes $T[0, \dots, n - l^* - 1]$ and T^β denotes $T[n - l^*, \dots, n - 1]$. Otherwise, the tag is computed without splitting (lower variant). The key used for the block cipher E is the same for every encryption. Hence, it can be precomputed. The returned tag is the n -bit value T . The $(n - l)$ -bit value Z is discarded. The decryption process works in a similar way from 'left to right', only the block cipher component E is replaced by its counterpart E^{-1} apart from one exception: the first call computes τ .

2.4 McOE-G

The third and last member of the McOE family presented in this paper is given by McOE-G. This version updates the chaining value by applying an almost XOR-universal hash function to the XOR result of the previous message block and ciphertext block (see Figure 4). In our practical implementation, we use the Galois-Field multiplication for H , *i.e.*, the key K_2 is multiplied with the chaining value over $\text{GF}(2^{128})$ defined by the low-weight irreducible polynomial $g(x) = x^{128} + x^7 + x^2 + x + 1$ as used in OCB [40] and GCM [34].

The tweakable block cipher \tilde{E}_K is then defined as follows

$$\tilde{E}_K(U, M) := E_{K_1}(M \oplus H_{K_2}(U)) \oplus H_{K_2}(U),$$

where E_K denotes a common block cipher, M the message, and U (chaining value) the tweak. The key K is denoted by the concatenation of K_1 and K_2 , *i.e.*, $K = K_1 || K_2$.

For McOE-G we also present a version with the *tag-splitting* approach, which was introduced before (see Section 6). Here, an additional key K_2 is required for the XOR-universal hash function H_{K_2} as shown in Figure 4. McOE-G can be easily extended to smoothly handle associated data, *i.e.*, data that is not encrypted but only authenticated. The security proofs considering associated data are given in Section 7.

2.5 Benchmarking

This section is about measuring the performance of the two presented members of the McOE family. The reference values are given by the CBC encryption scheme. Note, that the implementation of the CBC mode does not contain authentication. The results of our *naive* implementation based on common reference code are illustrated in Table 3.

Block cipher	Impl.	Message length in Bytes									
		64	128	256	512	1024	2048	4096	8192	16384	32768
McOE-X-AES	software	31.2	26.3	23.9	22.7	22	21.7	21.6	21.5	21.5	21.5
McOE-X-AES	AES-NI	14.2	12.2	11.2	10.7	10.5	10.4	10.4	10.3	10.3	10.3
McOE-X-Threefish	software	19.5	13.1	9.9	8.3	7.5	7.1	6.9	6.8	6.8	6.7
McOE-G-AES	software	33	27.9	25.4	24.1	23.5	23.2	23	22.9	22.8	22.8
McOE-G-AES	GF-NI/AES-NI	12.5	10.6	9.7	9.3	9	8.9	8.9	8.8	8.8	8.8
AES-CBC encryption	software	38.3	35.9	13.5	13.3	13.2	13.2	13.1	13.1	13.1	13.1
AES-CBC encryption	AES-NI	4	3.7	3.6	3.5	3.5	3.5	3.5	3.5	3.5	3.5

Table 3. Performance values (cycles-per-byte, single core), measured on a Core i5 540M for AES-128 and Threefish-512. McOE-X is the main contribution in the current paper, McOE-G is a variant using Galois field arithmetic. For comparison, we also provide the performance of unauthenticated AES-CBC. The AES software implementation is based on Gladman [16], whereas the hardware implementation is based on the Intel AES-NI Sample Library [12]. The Threefish implementation is based on the NIST/SHA-3 reference source as provided by the Skein authors [35]. Finally, the implementation of Galois-Field NI multiplication (GF-NI) is based on the example code from [19].

3 On-Line Authenticated Encryption and Related Notions

3.1 Definitions

Length of Longest Common Prefix (LLCP_n). The length of a string $x \in \{0, 1\}^n$ is denoted by $|x| := n$. For integers $n, \ell, d \geq 1$, set $D_n^d = (\{0, 1\}^n)^d$, and $D_n^* := \bigcup_{d \geq 0} D_n^d$, and $D_{\ell, n} = \bigcup_{0 \leq d \leq \ell} D_n^d$. Note that D_n^0 only contains the empty string. For $M \in D_n^d$, we write $M = (M_1, \dots, M_d)$ with $M_1, \dots, M_d \in D_n$. For $P, R \in D_n^*$, say, $P \in D_n^p$ and $R \in D_n^r$, we define the *length of the longest common n-prefix* of P and R as

$$\text{LLCP}_n(P, R) = \max_i \{P_1 = R_1, \dots, P_i = R_i\}.$$

For a non-empty set \mathcal{Q} of strings in D_n^* we define $\text{LLCP}_n(\mathcal{Q}, P)$ as $\max_{q \in \mathcal{Q}} \{\text{LLCP}_n(q, P)\}$. For example, if $P \in \mathcal{Q}$, then $\text{LLCP}_n(\mathcal{Q}, P) = |P|/n$.

For convenience, we introduce a notation for a *restriction on a set*. Let $\mathcal{Q} = \{0, 1\}^a \times \{0, 1\}^b \times \{0, 1\}^c$, then we denote by $\mathcal{Q}|_{b,c} = \{(B, C) \mid \exists A : (A, B, C) \in \mathcal{Q}\}$ as the restriction of \mathcal{Q} to B and C. This generalizes in the obvious way.

Game Based Proofs. Most of the proofs in this paper use the concept of *game-playing proofs*. The presented games in this paper are written in a language heavily based on \mathcal{L} , that was introduced by Bellare and Rogaway in [5]. A game has three kinds of functions: An initialization function `Initialize`, a finalization function `Finalize`, and oracle functions. Any adversary A that is playing a game calls at first the `Initialize` function. In the following, A makes some oracle queries and finally it ends the game by invoking `Finalize`. For adversaries, a function of a game is a black box. They have no access to any local or global variable of any game. An adversary wins the game if and only if `Finalize` returns `true`. We denote $\Pr[A^G \Rightarrow 1]$ as the probability that the adversary wins the game G .

Note, in this paper we usually use a three digit line number which follows the notation of Bellare and Rogaway where the first digit denotes the Game, *e.g.*, 444 denotes the 44-th line of Game G_4 .

3.2 Block Ciphers and On-Line Permutations

Block Cipher. A (k, n) -block cipher is a keyed family of permutations consisting of two paired algorithms $E : D_k \times D_n \rightarrow D_n$ and $E^{-1} : D_k \times D_n \rightarrow D_n$, accepting a k -bit key and an input from D_n for some $k, n > 0$. For $n > 0$, $Block(k, n)$ is the set of all (k, n) -block ciphers. For any $E \in Block(k, n)$ and fixed key $K \in D_k$, the decryption $E_K^{-1}(Y) := E^{-1}(K, Y)$ is the inverse function of encryption $E_K(X) := E(K, X)$, so that $E_K^{-1}(E_K(X)) = X$ holds for any $X \in D_n$.

We follow the usual convention to write oracles, that are provided to an algorithm, as superscripts. We define the related-key PRP-security of a block cipher E by the success probability of an adversary trying to distinguish between the block cipher and a random permutation.

Definition 3. Let $E \in Block(k, n)$ and let denote E^{-1} its corresponding inverse. Let $\varphi : D_k \times D_n \rightarrow D_k$. A fixed related key adversary A has access to an oracle E with two parameters such that she can query either $E_{\varphi(K, \cdot)}(\cdot)$ or $E_{\varphi(K, \cdot)}^{-1}(\cdot)$.

Let $PERM(n, n)$ be the set of n -bit permutations, such that the first parameter models the permutation and the second parameter the value that is to be permuted, i.e., for $\pi \in PERM(n, n)$ it holds that $\pi(Z, \cdot)$ is a random permutation for any given value of Z .

The related-key (RK) advantage [32] of A in breaking E is then defined by

$$\begin{aligned} \mathbf{Adv}_E^{\text{RK-CPA-PRP}}(A) &= |\Pr[K \xleftarrow{\$} D_k A^{E_{\varphi(K, \cdot)}(\cdot)} \Rightarrow 1] - \Pr[\pi \xleftarrow{\$} Perm(n, n) : A^{\pi(\cdot, \cdot)} \Rightarrow 1]| \\ \mathbf{Adv}_{E, E^{-1}}^{\text{RK-CCA-PRP}}(A) &= |\Pr[K \xleftarrow{\$} D_k : A^{E_{\varphi(K, \cdot)}(\cdot), E_{\varphi(K, \cdot)}^{-1}(\cdot)} \Rightarrow 1] \\ &\quad - \Pr[\pi \xleftarrow{\$} Perm(n, n) : A^{\pi(\cdot, \cdot), \pi^{-1}(\cdot, \cdot)} \Rightarrow 1]|. \end{aligned}$$

Tweakable Block Cipher. The concept of tweakable block ciphers was introduced by Liskov *et al.* in [31]. The design is based on a common block cipher, which is extended by a so-called tweak. A tweakable (k, v, n) block cipher is a family of functions consisting of two paired algorithms $\tilde{E} : D_k \times D_v \times D_n \rightarrow D_n$ and $\tilde{E}^{-1} : D_k \times D_v \times D_n \rightarrow D_n$, accepting a key $K \in D_k$, a tweak $V \in D_v$, and an input from D_n for some $k, v, n > 0$. $Block(k, v, n)$ is the set of all (k, v, n) -tweakable block ciphers. For any $\tilde{E} \in Block(k, v, n)$ and two fixed values $K \in D_k$ and $V \in D_v$, the decryption $\tilde{E}_K^{-1}(V, Y) := E^{-1}(K, V, Y)$ is the inverse function of the encryption $\tilde{E}_K(V, X) := \tilde{E}(K, V, X)$, i.e., $\tilde{E}_K(K, V, \cdot)$ is a permutation.

Definition 4. Let $\tilde{E} \in Block(k, v, n)$ and denote by \tilde{E}^{-1} the corresponding inverse. A fixed adversary A has access to an \tilde{E} oracle with three parameters such that she can query either $\tilde{E}_K(\cdot, \cdot)$ or $\tilde{E}_K^{-1}(\cdot, \cdot)$.

Let $Tperm(v, n)$ be the set of n -bit permutations such that the first parameter models the permutation and the second parameter the value that is to be permuted, i.e., for $\pi \in Tperm(v, n)$ it holds that $\pi(Z, \cdot)$ is a random permutation for any given value of Z .

The advantage for an adversary A to distinguish \tilde{E} from a randomly chosen permutation from $Tperm(v, n)$ is defined as

$$\begin{aligned} \mathbf{Adv}_E^{\text{T-IND-CPA}}(A) &= |\Pr[K \xleftarrow{\$} D_k : A^{\tilde{E}_K(\cdot, \cdot)} \Rightarrow 1] - \Pr[\pi \xleftarrow{\$} Tperm(v, n) : A^{\pi(\cdot, \cdot)} \Rightarrow 1]| \\ \mathbf{Adv}_{E, E^{-1}}^{\text{T-IND-CCA}}(A) &= |\Pr[K \xleftarrow{\$} D_k : A^{\tilde{E}_K(\cdot, \cdot), \tilde{E}_K^{-1}(\cdot, \cdot)} \Rightarrow 1] \\ &\quad - \Pr[\pi \xleftarrow{\$} Tperm(v, n) : A^{\pi(\cdot, \cdot), \pi^{-1}(\cdot, \cdot)} \Rightarrow 1]|. \end{aligned}$$

```

1 OPerm( $V, M$ )
2    $(j, p) \leftarrow \text{LLCP}_n^*(\mathcal{Q}_{|V, M}, (V, M));$ 
3    $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(V, M, C)\};$ 
4   for  $i = 1, \dots, p$  do
5      $C_i \leftarrow C_i^j;$ 
6   for  $i = p + 1, \dots, |M|/n$  do
7      $C_i \xleftarrow{\$} D_n \setminus D[V||M_1||\dots||M_{i-1}];$ 
8      $D[V||M_1||\dots||M_{i-1}] \leftarrow D[V||M_1||\dots||M_{i-1}] \cup C_i;$ 
9   return  $C;$ 

```

Fig. 5. Lazy-sampling implementation of a stateful (n -)on-line permutation. In Line 2, the Oracle LLCP_n^* is invoked returning (j, p) where p denotes the length of the prefix n determined via LLCP_n and j denotes the index in $\mathcal{Q}_{|V, M}$. In line 5, we denote by C_i^j the i -th n -bit block of the j -th entry in $\mathcal{Q}_{|C}$.

On-Line Permutation (OPerm). We aim for larger permutations that not only permute single but multiple-block messages. We say a permutation $P : D_n^* \rightarrow D_n^*$ is (n -)on-line iff the i -th output block is determined completely by the first i blocks of the input. Let denote $\text{OPerm}_{n,*}$ the set of all D_n^* on-line permutations. It is easy to extend the definition with a state space D_v . Let $\text{OPerm}_{n,*}^v$ denote the set of all functions from $D_v \times D_n^*$ to D_n^* . Then for each $f \in \text{OPerm}_{n,*}^v$ and $V \in D_v$, the function $f(V, \cdot)$ is an (n -)on-line permutation. Figure 5 illustrates a lazy sampling implementation of $\text{OPerm}_{n,*}^v$.

Next, we introduce the formal definition of a family of (n -)on-line functions which is the basic design principle of the MCOE family.

Definition 5. Let $n, k, v \geq 0$, $K \in D_k$, $V \in D_v$. A family of functions $F : D_k \times D_v \times D_n^* \rightarrow D_n^*$ is (n -)on-line if for any instance of this family determined by K, V , $F(K, V, \cdot)$ is a permutation and for any message $M = (M_1, M_2, \dots, M_\ell)$ there exists a family of functions $f^i : D_k \times (D_v \times D_n^{i-1}) \times D_n \rightarrow D_n$, $i = 1, \dots, \ell$ such that

$$\begin{aligned}
\Pi(K, V, M) = & f_K^1((V, \emptyset), M_1) || f_K^2((V, M_1), M_2) \\
& || \dots || \\
& f_K^{\ell-1}((V, M_1 || \dots || M_{\ell-2}), M_{\ell-1}) || f_K^\ell((V, M_1 || \dots || M_{\ell-1}), M_\ell),
\end{aligned}$$

where “||” is the concatenation of strings.

An encryption scheme is (n -)on-line if the encryption function is (n -)on-line. A thorough discussion of on-line encryption and its properties can be found in [1].

Proposition 1. Let F be an (n -)on-line function as defined in Definition 5. Then, all $f_K^i(V, \cdot)$ are n -bit permutations.

The proof is similar to Proposition 3.4 of [1].

Let F be a family of (n -)on-line functions. Assume that for each uniform randomly chosen F_K from F , each $f_K^i(V, \cdot)$ is a PRP. Then, it is easy to see that F_K is indistinguishable from the OPerm oracle as shown in Figure 5. We call such a family of (n -)on-line functions *on-line pseudo random permutations* (OPRP).

3.3 Authenticated Encryption (With Associated Data)

An authenticated-encryption scheme is a tuple $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ whose aim is to provide privacy and data integrity. The key generation function \mathcal{K} takes no input and returns a randomly chosen key K from the key space, e.g., from D_k . The encryption algorithm \mathcal{E} and the decryption algorithm \mathcal{D} are deterministic algorithms that map values from $D_k \times D_n^+ \times D_n^*$ to a string or – if the input is invalid – the value \perp . The header H consists of either only the initial value/nonce $V \in D_n$ (if no data is to be authenticated/checked in the encryption/decryption process) or is a combination

of V and a value from D_n^* : $H \in D_n^+$. For sake of convenience, we usually write $\mathcal{E}_K^H(M)$ for $\mathcal{E}(K, H, M)$ and $\mathcal{D}_K^H(M)$ for $\mathcal{D}(K, H, M)$, where the message M is chosen from D_n^* , $H \in D_n^+$, and a key from the key space. We require $\mathcal{D}_K^H(\mathcal{E}_K^H(M)) = M$ for any possible K, M, H , and define the tag size for a message $M \in D_n^*$ and header $H \in D_n^+$ as $\text{TAG}(H, M) := |\mathcal{E}_K^H(M)| - |M|$. We denote an authenticated-encryption scheme with the requirement that the initial vector V is only used once in a *nonce-based* scheme. Otherwise, we call it *deterministic*. Similarly, we call an adversary *nonce-respecting* (NR) if no nonce is used twice for any query. Otherwise, the adversary is called *nonce-ignoring* (NI).

Security

Game $G_{\text{CPA}}, G_{\text{CCA3}}$ 1 Initialize (ω, ν) 2 $b \xleftarrow{\$} \{0, 1\}$; 3 if ($b=1$) then 4 $K \leftarrow \mathcal{K}()$; 5 Finalize (d) 6 return ($b = d$);	10 Encrypt (H, M) 11 if ($\nu = \text{NR}$ and $V \in B$) then 12 return \perp ; 13 if ($b=1$) then 14 $C \leftarrow \mathcal{E}_K(H, M)$; 15 else 16 $C \leftarrow \mathcal{E}^\omega(H, M)$; 17 $B \leftarrow B \cup \{V\}$; 18 $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, C)\}$; 19 return C ;	20 Decrypt (H, C) 21 if ($(H, C) \in \mathcal{Q}$) then 22 return \perp ; 23 if ($b=1$) then 24 $M \leftarrow \mathcal{D}_K(H, C)$; 25 else 26 $M \leftarrow \perp(H, C)$; 27 return M ;
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Fig. 6. $G_{\text{CPA}}(\omega, \nu)$ is the $\text{CPA}_H^{(\omega, \nu)}$ -Game and $G_{\text{CCA3}}(\omega, \nu)$ the $\text{CCA3}_H^{(\omega, \nu)}$ -Game where $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$. Game G_{CCA3} contains the code in the box while G_{CPA} does not. The oracle $\mathcal{E}^{\text{AE}}(H, M)$ returns a string of length $|M| + \text{TAG}(H, M)$. V denotes the last block of the header representing the nonce/initial value.

4 Security Notions for On-Line Authenticated Encryption

The security notions and their relations for nonce-based authenticated-encryption schemes are introduced in [3,4,27,37,40]. Furthermore, Rogaway and Shrimpton introduced in [42] notions for deterministic authenticated-encryption schemes. In order to have one convenient toolset of notions, we adopt the notion of CCA3-security suggested in [42] as a *natural strengthening* of CCA2-security.

We parameterize our definition in order to define different – but closely related – notions by explicitly stating whether we mean an on-line or off-line scheme, $\omega \in \{\text{AE}, \text{OAE}\}$, and stating the adversary behavior as either nonce-respecting or nonce-ignoring, $\nu \in \{\text{NR}, \text{NI}\}$.

Definition 6 (CCA3(ω, ν)). Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be an authenticated encryption scheme with header space D_n^+ and message space D_n^* , and fix an adversary A . The advantage of A to break Π is defined as

$$\mathbf{Adv}_{\Pi}^{\text{CCA3}(\omega, \nu)}(A) = \left| \Pr \left[K \xleftarrow{\$} \mathcal{K} : A^{\mathcal{E}_K(\cdot, \cdot), \mathcal{D}_K(\cdot, \cdot)} \Rightarrow 1 \right] - \Pr \left[A^{\mathcal{E}^\omega(\cdot, \cdot), \perp(\cdot, \cdot)} \Rightarrow 1 \right] \right|.$$

The adversary's random-bits oracle, $\mathcal{E}^{\text{AE}}(\cdot, \cdot)$ or $\mathcal{E}^{\text{OAE}}(\cdot, \cdot)$, returns on a query with header $H \in D_n^+$ and plaintext $X \in D_n^*$ a random string of length $|\mathcal{E}_K(M)|$ which is either on-line or not, depending on the variable ω . The oracle $\perp(\cdot, \cdot)$ returns \perp on every input. Wlog, we assume that the adversary A never asks a query which the answer is already known. It is easy to see that we can rewrite the term given in Definition 6 by

$$\mathbf{Adv}_{\Pi}^{\text{CCA3}(\omega, \nu)}(A) = \left| \Pr \left[K \xleftarrow{\$} \mathcal{K} : A^{\mathcal{E}_K(\cdot, \cdot), \mathcal{D}_K(\cdot, \cdot)} \Rightarrow 1 \right] - \Pr \left[K \xleftarrow{\$} \mathcal{K} : A^{\mathcal{E}_K(\cdot, \cdot), \perp(\cdot, \cdot)} \Rightarrow 1 \right] \right| \quad (1)$$

$$+ \Pr \left[K \xleftarrow{\$} \mathcal{K} : A^{\mathcal{E}_K(\cdot, \cdot), \perp(\cdot, \cdot)} \Rightarrow 1 \right] - \Pr \left[A^{\mathcal{E}^\omega(\cdot, \cdot), \perp(\cdot, \cdot)} \Rightarrow 1 \right] \Big|. \quad (2)$$

One can interpret (1) as the advantage that an adversary can generate an existential forgery (2) as the advantage that an CPA adversary can distinguish the ciphertext from a random bit string. Using this decomposition as a starting point, we now define ciphertext integrity and what we mean by a CPA adversary on authenticated-encryption schemes. From now on, our definitions are based on the game playing methodology. For example, we can restate Definition 6 using the game G_{CCA3} given in Figure 6 as

$$\mathbf{Adv}_{\Pi}^{\text{CCA3}(\omega, \nu)}(A) = 2 \left| \Pr[A^{G_{\text{CCA3}}(\omega, \nu)} \Rightarrow 1] - 0.5 \right|.$$

We denote by $\mathbf{Adv}_{\Pi}^{\text{CCA3}(\omega, \nu)}(q, \ell, t)$ the maximum advantage over all $\text{CCA3}(\omega, \nu)$ -adversaries run in time at most t , ask a total maximum of q queries to \mathcal{E} and \mathcal{D} , and query at most ℓ blocks.

Game $G_{\text{INT-CTXT}}$ 1 Initialize (ν) 2 $K \leftarrow \mathcal{K}()$; 3 Finalize () 4 return win ;	10 Encrypt (H, M) 11 if ($\nu = \text{NR}$ and $V \in B$) then 12 return \perp ; 13 $C \leftarrow \mathcal{E}_K(H, M)$; 14 $B \leftarrow B \cup \{V\}$; 15 $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, C)\}$; 16 return C ;	20 Verify (H, C) 21 $M \leftarrow \mathcal{D}_K(H, C)$; 22 if ($(H, C) \notin \mathcal{Q}$ and $M \neq \perp$) then 23 $\text{win} \leftarrow \text{true}$; 24 return ($M \neq \perp$) ;
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Fig. 7. Game $G_{\text{INT-CTXT}}(\nu)$ is the $\text{INT-CTXT}_{\Pi}^{\omega, \nu}$ game where $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$. V denotes the last block of the header representing the nonce/initial value.

4.1 Privacy and Integrity Notions for Authenticated Encryption Schemes.

Similarly, we define the privacy and integrity of an (on-line) authenticated encryption scheme $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ with header space D_n^+ , message space D_n^* , and tag-size function $\text{TAG}(H, M)$ as follows.

Definition 7. Let $G_{\text{CPA}}(\omega, \nu)$ be the $\text{CPA}_{\Pi}^{\omega, \nu}$ game given in Figure 6. Fix an adversary A . Then, the advantage of A breaking Π is defined as

$$\text{Adv}_{\Pi}^{\text{CPA}(\omega, \nu)}(A) \leq 2 \left| \Pr[A^{G_{\text{CPA}}(\omega, \nu)} \Rightarrow 1] - 0.5 \right|.$$

Definition 8. Let $G_{\text{INT-CTXT}}(\nu)$ be the $\text{INT-CTXT}_{\Pi}^{\nu}$ -game given in Figure 7. Fix an adversary A . The advantage of A breaking Π is defined as

$$\text{Adv}_{\Pi}^{\text{INT-CTXT}(\nu)}(A) \leq \Pr[A^{G_{\text{INT-CTXT}}(\nu)} \Rightarrow 1].$$

We denote $\text{Adv}_{\Pi}^{\text{CPA}(\omega, \nu)}(q, t, \ell)$ and $\text{Adv}_{\Pi}^{\text{INT-CTXT}(\nu)}(q, t, \ell)$ as the maximum advantage over all $\text{CPA}(\omega, \nu)$ resp. $\text{INT-CTXT}(\nu)$ adversaries that run in time at most t , ask a total maximum of q queries to \mathcal{E} and \mathcal{D} , and query at most ℓ blocks.

4.2 CCA3 is equal to INT-CTXT plus CPA-security.

Now we generalize Theorem 3.2 from Bellare and Namprepre [3]. It simply states the equivalence of a scheme which is CCA3 secure and both INT-CTXT and CPA secure (often denoted as IND-CPA secure). These statements hold in the on-line and off-line case.

Theorem 1. Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be an authenticated-encryption scheme. Fix $\omega \in \{\text{AE}, \text{OAE}\}$ and $\nu \in \{\text{NR}, \text{NI}\}$. Let A be an $\text{CCA3}(\omega, \nu)_{\Pi}$ -adversary running in time at most t , making at most q queries with a total length of at most ℓ blocks. Then, there exists a $\text{CPA}(\omega, \nu)$ -adversary A_p and an $\text{INT-CTXT}(\omega, \nu)$ -adversary A_c such that

$$\text{Adv}_{\Pi}^{\text{CCA3}(\omega, \nu)}(A) \leq \text{Adv}_{\Pi}^{\text{CPA}(\omega, \nu)}(A_p) + \text{Adv}_{\Pi}^{\text{INT-CTXT}(\omega, \nu)}(A_c).$$

Furthermore, A_c and A_p run in time $O(t)$ and both make at most q queries in each case.

The proof is given in Appendix B.

4.3 Relations between PRP and OPRP.

Let us proceed from on-line authenticated-encryption schemes in a common case, where we consider an adversary A to be nonce-respecting and the regarded scheme to be $\text{CCA3}^{\text{AE}, \text{NR}}$ secure. For such schemes, it is desirable to obtain a certain level of security, even in a misuse-scenario. Due to the nature of this scenario, an on-line authenticated encryption scheme can *solely* be $\text{CCA3}^{\text{OAE}, \text{NI}}$ -secure. Hence, it is of great interest to determine the relation between both security notions.

Lemma 1. *Let F be an OPRP, introduced in Section 3.2, and A be a nonce-respecting adversary. Then*

$$\mathbf{Adv}_F^{PRP^{(nr)}}(A) = |\Pr[K \stackrel{\$}{\leftarrow} \mathcal{K} : A^{F_K} \Rightarrow 1] - \Pr[p \stackrel{\$}{\leftarrow} PRP : A^p \Rightarrow 1]| = 0.$$

Proof (Sketch). Let (V_i, M_i) denote the i -th encryption query. For each V_i, V_j , with $i \neq j$, it holds true that $V_i \neq V_j$ since we assume a nonce-respecting adversary. Consequently, $F_K(V_i, \cdot)$ and $F_K(V_j, \cdot)$ are two independent PRPs due to the fact that all $f_k^\ell(V||X, \cdot)$ with $X \in D_n^{\ell-1}$ are PRPs. This implies that $F_K(V, \cdot)$ is a PRP. \square

This Lemma shows that a $\text{CCA3}^{\text{OAE,NI}}$ -secure on-line authenticated encryption scheme is also $\text{CCA3}^{\text{AE,NR}}$ -secure against nonce-respecting adversaries.

5 Security of the Generic McOE Scheme

In this section, we analyze the security of the generic McOE scheme, which was introduced by Definition 1 and 2 (cf. Section 2). We show that McOE achieves our two-fold goal, by proving that it guarantees a certain minimum, well-defined security against a nonce-ignoring adversary. More formally, we show that McOE is $\text{CCA3}^{\text{OAE,NI}}$ secure, which implies that McOE is also fully secure against a nonce-respecting adversary, *i.e.*, $\text{CCA3}^{\text{AE,NR}}$ -secure (cf. Section 4.3).

5.1 Security Analysis of McOE without Tag-Splitting

Now we proceed to show the security of McOE. Therefore, this we use the results of Theorem 1 and show the INT-CTXT- and RK-CPA-PRP-security separately.

Theorem 2. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a McOE scheme as in Definition 1, *i.e.*, \mathcal{K} is the key-derivation function, $\mathcal{E} = \mathbf{EncryptAuthenticate}$, and $\mathcal{D} = \mathbf{DecryptAuthenticate}$. Then*

$$\mathbf{Adv}_{\Pi}^{\text{CCA3}(\text{OAE,NI})}(q, \ell, t) \leq \frac{3(q + \ell)(q + \ell + 1) + 4q + 3\ell}{2^n - (q + \ell)} + 3\mathbf{Adv}_{\tilde{E}, \tilde{E}^{-1}}^{\text{T-IND-CCA}}(q + \ell, O(t)).$$

Proof. The proof follows from Theorem 1 together with Lemmas 2 and 3. \square

For the sake of simplification, we provide an upper bound which is much easier to grasp than the original bound, but not as tight as the original bound given in the theorem above.

Corollary 1. *Assume that $\ell \geq 7$, $\ell \geq q$ and the T-IND-CCA-advantage is at most δ for an adversary whose amount of queries is at most $q + \ell$ and its running time is $O(t)$. Then, the following bound holds:*

$$\mathbf{Adv}_{\Pi}^{\text{CCA3}(\text{OAE,NI})}(q, \ell, t) \leq \frac{14\ell^2}{2^n - (2\ell)} + \delta.$$

Lemma 2. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a McOE scheme as in Definition 1. Let q be the number of total queries an adversary A is allowed to ask and ℓ be an integer representing the total length in blocks of the queries to \mathcal{E} and \mathcal{D} . Then,*

$$\mathbf{Adv}_{\Pi}^{\text{INT-CTXT}(\text{NI})}(q, \ell, t) \leq \frac{(q + \ell)(q + \ell + 1)}{2^n - (q + \ell)} + \frac{2q + \ell}{2^n - (q + \ell)} + \mathbf{Adv}_{\tilde{E}, \tilde{E}^{-1}}^{\text{T-IND-CCA}}(q + \ell).$$

```

1 Initialize ()
2    $K \xleftarrow{\$} \mathcal{K}$ ;
3    $B_1 \leftarrow \{0^n\}$ ;

100 Encrypt( $H, M$ ) Game  $G_1$ 
101    $L_H \leftarrow |H|/n$ ;  $L \leftarrow |M|/n$ ;
102    $U \leftarrow 0^n$ ;
103   for  $i = 1, \dots, L_H$  do
104      $\tau \leftarrow \tilde{E}_K(U, H_i)$ ;
105      $U \leftarrow H_i \oplus \tau$ ;
106   for  $i = 1, \dots, L$  do
107      $C_i \leftarrow \tilde{E}_K(U, M_i)$ ;
108      $U \leftarrow C_i \oplus M_i$ ;
109    $T \leftarrow \tilde{E}_K(U, \tau)$ ;
110    $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, M, C, T)\}$ ;
111   return  $(C_1, \dots, C_L, T)$ ;

200 Encrypt( $H, M$ ) Game  $G_2, \boxed{G_3}$ 
201    $L_H \leftarrow |H|/n$ ;  $L \leftarrow |M|/n$ ;
202    $B_2 \leftarrow B_1 \cup \{H\}$ ;
203    $p \leftarrow \text{LLCP}_n(\mathcal{Q}_{|H, M}, (H, M))$ ;
204    $U \leftarrow 0^n$ ;
205   for  $i = 1, \dots, L_H$  do
206      $\tau \leftarrow \tilde{E}_K(U, H_i)$ ;
207      $U \leftarrow H_i \oplus \tau$ ;
208     if  $(U \in B_1 \text{ and } i > p)$  then
209       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B_1$ ;
210      $B_1 \leftarrow B_1 \cup \{U\}$ ;
211   for  $i = 1, \dots, L$  do
212      $C_i \leftarrow \tilde{E}_K(U, M_i)$ ;
213      $U \leftarrow C_i \oplus M_i$ ;
214     if  $(U \in B_1 \text{ and } i + L_H > p)$  then
215       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B_1$ ;
216      $B_1 \leftarrow B_1 \cup \{U\}$ ;
217    $T \leftarrow \tilde{E}_K(U, \tau)$ ;
218    $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, M, C, T)\}$ ;
219   return  $(C_1, \dots, C_L, T)$ ;

4 Finalize ()
5   return win;

112 Verify( $H, C, T$ ) Game  $G_1$ 
113    $L_H \leftarrow |H|/n$ ;  $L \leftarrow |C|/n$ ;
114    $U \leftarrow 0^n$ ;
115   for  $i = 1, \dots, L_H$  do
116      $\tau \leftarrow \tilde{E}_K(U, H_i)$ ;
117      $U \leftarrow H_i \oplus \tau$ ;
118   for  $i = 1, \dots, L$  do
119      $M_i \leftarrow \tilde{E}_K^{-1}(U, C_i)$ ;
120      $U \leftarrow C_i \oplus M_i$ ;
121   if  $(T = \tilde{E}_K(U, \tau) \text{ and } (H, C) \notin \mathcal{Q}_{|H, C})$  then
122     win  $\leftarrow$  true;
123    $\mathcal{Q} \leftarrow \mathcal{Q} \cup (H, \perp, C, \perp)$ ;
124   return  $(T = \tilde{E}_K(U, \tau))$ ;

220 Verify( $H, C, T$ ) Game  $G_2, \boxed{G_3}$ 
221    $L_H \leftarrow |H|/n$ ;  $L \leftarrow |C|/n$ ;
222    $p \leftarrow \text{LLCP}_n(\mathcal{Q}_{|H, M}, (H, M))$ ;
223    $U \leftarrow 0^n$ ;
224   for  $i = 1, \dots, L_H$  do
225      $\tau \leftarrow \tilde{E}_K(U, H_i)$ ;
226      $U \leftarrow H_i \oplus \tau$ ;
227     if  $(U \in B_1 \text{ and } i > p)$  then
228       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B_1$ ;
229      $B_1 \leftarrow B_1 \cup \{U\}$ ;
230   for  $i = 1, \dots, L - 1$  do
231      $M_i \leftarrow \tilde{E}_K^{-1}(U, C_i)$ ;
232      $U \leftarrow C_i \oplus M_i$ ;
233     if  $(U \in B_1 \text{ and } i + L_H > p)$  then
234       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B_1$ ;
235      $B_1 \leftarrow B_1 \cup \{U\}$ ;
236    $M_L \leftarrow \tilde{E}_K^{-1}(U, C_L)$ ;
237    $U \leftarrow C_L \oplus M_L$ ;
238   if  $(U \in B_1 \text{ and } H \notin B_2)$  then
239     bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B_1$ ;
240   if  $(T = \tilde{E}_K(U, \tau) \text{ and } (H, C, T) \notin \mathcal{Q}_{|H, C, T})$  then
241     win  $\leftarrow$  true;
242    $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, \perp, C, \perp)\}$ ;
243    $B \leftarrow B \cup \{U\}$ ;
244   return  $(T = \tilde{E}_K(U, \tau))$ ;

```

Fig. 8. Games G_1 - G_3 for the proof of Lemma 2. Game G_3 contains the code in the box while G_2 does not.

Proof. Our bound is derived from game-playing arguments. Consider games G_1 - G_3 of Figure 8 and a fixed adversary A asking at most q queries with a total length of at most ℓ blocks. The functions **Initialize** and **Finalize** are identical for all games in this proof. Let denote G_0 as the Game INT-CTXT(NI) as defined in Figure 7. Definition 8 states that

$$\mathbf{Adv}_H^{\text{INT-CTXT(NI)}}(A) \leq \Pr[A^{G_0} \Rightarrow 1].$$

In G_1 , the encryption and verify placeholders are replaced by their generic MCOE counterparts from Definition 1. Clearly, $\Pr[A^{G_0} \Rightarrow 1] = \Pr[A^{G_1} \Rightarrow 1]$. In the following, we discuss the differences between G_1 and G_2 . The set B_1 is initialized with 0^n and then collects all new key-input values U , which are computed during the encryption or verification process (in lines 204, 207, 213, 223, 226, 232 and 237).

In lines 203 and 222, the LLCP_n oracle is inquired. Finally, the variable **bad** is set to **true** if one of the if-conditions in lines 208, 214, 227, 233, or 238 is **true**. *None* of these modifications affect the values returned to the adversary and therefore

$$\Pr[A^{G_1} \Rightarrow 1] = \Pr[A^{G_2} \Rightarrow 1].$$

For our further discussion we require another game G_4 which is explained in more detail later in this proof². It follows that

$$\begin{aligned} \Pr[A^{G_2} \Rightarrow 1] &= \Pr[A^{G_3} \Rightarrow 1] + |\Pr[A^{G_2} \Rightarrow 1] - \Pr[A^{G_3} \Rightarrow 1]| \\ &\leq \Pr[A^{G_3} \Rightarrow 1] + \Pr[A^{G_3} \text{ sets bad}] \\ &\leq \Pr[A^{G_4} \Rightarrow 1] + |\Pr[A^{G_3} \Rightarrow 1] - \Pr[A^{G_4} \Rightarrow 1]| + \Pr[A^{G_3} \text{ sets bad}]. \end{aligned} \quad (3)$$

Now we proceed to upper bound the three terms contained in (3) – in right-to-left order. The success probability of game G_3 does not differ from the success probability of G_2 unless a chaining value U occurs twice. In this case, the adversary must (1) either have ‘found’ a collision for $\tilde{E}_K(X, Y) \oplus Y$, *i.e.*, she stumbles over (X, Y) and (X', Y') such that $\tilde{E}_K(X, Y) \oplus Y = \tilde{E}_K(X', Y') \oplus Y'$, or (2), must have found a preimage of 0^n , which is always the starting point of our chain. Note, the value 0^n is initially stored in the set B_1 . In both cases, the variable **bad** would have been set to **true**, and it follows by [9] that

$$\Pr[A^{G_3} \text{ sets bad}] \leq \frac{(q + \ell)(q + \ell + 1)}{2^n - (q + \ell)} + \frac{q + \ell}{2^n - (q + \ell)}.$$

Next, we describe the new game G_4 . It is equal to G_3 *except* that the tweakable block cipher \tilde{E} and its inverse \tilde{E}^{-1} are replaced by the functions **EncryptBlock** and **DecryptBlock**, which are modeled as a set of pseudo random permutations, where the index is given by the tweak. We assume that they are implemented via lazy sampling. More precisely, the call $\tilde{E}_K(X, Y)$ is replaced by an invocation of **EncryptBlock** $_K(X, Y)$ and the call $\tilde{E}_K^{-1}(X, Y)$ is replaced by an invocation of **DecryptBlock** $_K(X, Y)$. We now upper bound the difference between G_3 and G_4 .

So, by definition of G_4 , we have

$$|\Pr[A^{G_3} \Rightarrow 1] - \Pr[A^{G_4} \Rightarrow 1]| \leq \mathbf{Adv}_{\tilde{E}, \tilde{E}^{-1}}^{\text{T-IND-CCA}}(q + \ell, O(t)).$$

Finally, we have to upper bound the advantage for the adversary A to win the game G_4 . A can only win this game if the condition in Line 238 (resp. 438 for game G_4) is **true**. As usual, we assume *wlog.* that A does not ask a question if the answer is already known which implies that $(H, C, T) \notin \mathcal{Q}_{H, C, T}$. For our analysis, we distinguish between three cases. There, we formally adjust Line 240 (*i.e.*, choose as the tag computation operation either \tilde{E} or \tilde{E}^{-1}) such that we always have enough randomness left for our result.

² Since the difference is very minor, we do not provide an extra figure.

Case 1: $H \in B_2$ and $U \in B$.

Since we already have computed τ in the past, the chance of success is at most by the probability $\Pr[\tilde{E}_K^{-1}(U, T) = \tau]$ which can be upper bounded by $1/(2^n - (q + \ell))$.

Case 2: $H \notin B_2$, and $U \notin B_1$.

Then, the tagging operation uses a new tweak and therefore, the output of $\tilde{E}_K(U, \tau)$ is uniformly distributed and the success probability is $\leq 1/2^n$.

Case 3: $H \in B_2$ and $U \notin B_1$.

In this case, the chance of success is at most $\Pr[\tilde{E}_K^{-1}(U, T) = \tau]$ which can be upper bounded by $1/2^n$.

Note, the 'missing' fourth case has been explicitly excluded by Line 240 (resp. 440). Since these three cases are mutually exclusive, we can upper bound the success probability for q queries by

$$\Pr[A^{G_4} \Rightarrow 1] \leq \frac{q}{2^n - (q + \ell)}.$$

Our claim follows by adding up the individual bounds. □

Lemma 3. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a MCOE scheme as in Definition 1 (i). Let q be the number of total queries an adversary A is allowed to ask and ℓ be an integer representing the total length of the queries to \mathcal{E} and \mathcal{D} . Then,*

$$\mathbf{Adv}_{\Pi}^{\text{CPA}(\text{AOE}, \text{NI})}(q, \ell, t) \leq 2 \left(\frac{(q + \ell)(q + \ell + 1)}{2^n - (q + \ell)} + \frac{q + \ell}{2^n - (q + \ell)} + \mathbf{Adv}_{\tilde{E}}^{\text{T-IND-CPA}}(q + \ell) \right).$$

Proof. Our bound is derived by game-playing arguments. Consider games G_1 and G_2 of Figure 11. The functions **Initialize** and **Finalize** are identical for any of them.

First we investigate the differences between the $\text{CPA}(\text{AOE}, \text{NI})$ -game from Figure 6 and G_1 from Figure 11. In G_1 we have replaced \mathcal{E} by its definition of MCOE, and $\w by an on-line encryption oracle **OnlinePermutation** (Line 102) that just models a 'perfect' OPRP, *i.e.*, for two plaintexts with an equal prefix it returns two ciphertexts that also share a prefix of the same length. Again we assume this oracle to be implemented by lazy sampling. Then, set B collects all chaining values (lines 113 and 119) in order to intercept the occurrence of two identical chaining values which would lead to two identical tweaks for the encryption of a block.

In Line 105, the oracle LLCP_n is invoked returning the length of the longest common prefix of (H, M) and $\mathcal{Q}_{|H, M}$.

Finally, the variable **bad** is set to **true** if (one of) the conditions of lines 111/211 or 117/217 holds. These changes do not affect the success probability of an adversary because the output of the oracle remains unchanged. More precisely, the distribution of the output does not change. This means that

$$\mathbf{Adv}_{\Pi}^{\text{CPA}(\text{AOE}, \text{NI})}(A) = 2 \cdot |\Pr[A^{G_1} \Rightarrow 1] - 0.5|,$$

and therefore, by common game-playing arguments – using a new game G_3 described shortly –,

$$\begin{aligned} \Pr[A^{G_1} \Rightarrow 1] &\leq \Pr[A^{G_2} \Rightarrow 1] + |\Pr[A^{G_1} \Rightarrow 1] - \Pr[A^{G_2} \Rightarrow 1]| \\ &\leq \Pr[A^{G_2} \Rightarrow 1] + \Pr[A^{G_2} \text{ sets bad}] \\ &\leq \Pr[A^{G_3} \Rightarrow 1] + |\Pr[A^{G_2} \Rightarrow 1] - \Pr[A^{G_3} \Rightarrow 1]| + \Pr[A^{G_2} \text{ sets bad}]. \end{aligned}$$

The success probability of game G_2 does not differ from the success probability of G_1 unless a chaining value U occurs twice. In this case, the adversary must either have found a collision for $\tilde{E}_K(X, Y) \oplus Y$, *i.e.*, she has found (X, Y) and (X', Y') such that $\tilde{E}_K(X, Y) \oplus Y = \tilde{E}_K(X', Y') \oplus Y'$

<pre> 1 Initialize () 2 $b \xleftarrow{\\$} \{0, 1\}; K \xleftarrow{\\$} \mathcal{K}(); B \leftarrow 0^n;$ </pre> <hr/> <pre> 100 Encrypt(H, M) Game $G_1, \boxed{G_2}$ 101 if ($b = 0$) then 102 $C \leftarrow \mathbf{OnlinePermutation}(H, M);$ 103 else 104 $L_H \leftarrow H /n; L \leftarrow M /n;$ 105 $p \leftarrow \text{LLCP}_n(\mathcal{Q}, (H, M));$ 106 $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, M)\};$ 107 $U \leftarrow 0^n;$ 108 for $i = 1, \dots, L_H$ do 109 $\tau \leftarrow \tilde{E}_K(U, H_i);$ 110 $U \leftarrow H_i \oplus \tau;$ </pre>	<pre> 3 Finalize(d) 4 return ($b=d$); </pre> <hr/> <pre> 111 if ($U \in B$ and $i > p$) then 112 bad \leftarrow true; $U \xleftarrow{\\$} \{0, 1\}^n \setminus B;$ 113 $B \leftarrow B \cup \{U\};$ 114 for $i = 1, \dots, L$ do 115 $C_i \leftarrow \tilde{E}_K(U, M_i);$ 116 $U \leftarrow C_i \oplus M_i;$ 117 if ($U \in B$ and $i + L_H > p$) then 118 bad \leftarrow true; $U \xleftarrow{\\$} \{0, 1\}^n \setminus B;$ 119 $B \leftarrow B \cup \{U\};$ 120 return $C;$ </pre>
--	--

Fig. 9. Games G_1 and G_2 for the proof of Lemma 3. Game G_2 contains the code in the box while G_1 does not.

or must have found a preimage of 0^n . In both cases, the variable **bad** would have been set to **true**, and it follows again by [9] that

$$\Pr[A^{G_2} \text{ sets bad}] \leq \frac{(q + \ell)(q + \ell + 1)}{2^n - (q + \ell)} + \frac{q + \ell}{2^n - (q + \ell)}.$$

The aforementioned new game G_3 is equal to the game G_2 *except* that the tweakable block cipher \tilde{E} and its inverse \tilde{E}^{-1} are replaced by the functions **EncryptBlock** and **DecryptBlock**, which are modeled as set of pseudo random permutations, where the index is given by the tweak. We assume that they are implemented via lazy sampling. More precisely, the call of $\tilde{E}_K(X, Y)$ is replaced by an invocation of **EncryptBlock** $_K(X, Y)$ and the call of $\tilde{E}_K^{-1}(X, Y)$ is replaced by an invocation of **DecryptBlock** $_K(X, Y)$. Now we upper bound the difference between G_2 and G_3 . So, by definition of G_4 , we have

$$|\Pr[A^{G_2} \Rightarrow 1] - \Pr[A^{G_3} \Rightarrow 1]| \leq \mathbf{Adv}_{\tilde{E}}^{\text{T-IND-CPA}}(q + \ell, O(t)).$$

Finally, we have to upper bound the advantage of an adversary A to win the game G_3 . Since U cannot collide and it is not possible to compute a preimage for any query, the algorithm for $b = 0$ is an OPRP, and therefore, the success probability to win G_3 for any adversary is 0.5, *i.e.*, she has no advantage in winning this game.

Our claim follows by adding up the individual bounds. □

5.2 Security Analysis of McOE with Tag-Splitting

Theorem 3. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a McOE scheme as in Definition 2, *i.e.*, \mathcal{K} is the key generation function, $\mathcal{E} = \mathbf{EAST}$ and $\mathcal{D} = \mathbf{DAST}$. For $q \leq 2^{n/2-2}$ we have*

$$\begin{aligned} \mathbf{Adv}_{\Pi}^{\text{CCA3(OAE,NI)}}(q, \ell, t) &\leq \frac{4(q + \ell + 2)(q + \ell + 3) + 6(2q + \ell)}{2^n - (q + \ell)} + \frac{3q(q + 1)}{2^n - q} \\ &\quad + \frac{q}{2^{n/2} - q} + 3\mathbf{Adv}_{E, E^{-1}}^{\text{T-IND-CCA}}(2q + \ell, O(t)). \end{aligned}$$

Proof. The proof follows from Theorem 1 together with Lemmas 4 and 6. □

For the sake of simplification we provide an upper bound which is much easier to grasp than the original bound, but not as tight as the original bound given in the theorem above.

Corollary 2. *Let assume that $\ell \geq 35$, $\ell \geq q$ and the T-IND-CCA-advantage is at most δ for an adversary which amount of queries is at most $q + \ell$ and its running time is $O(t)$. Then the following bound holds*

$$\mathbf{Adv}_{\Pi}^{\text{CCA3(OAE,NI)}}(q, \ell, t) \leq \frac{21\ell^2 + \ell}{2^n - (2\ell)} + \frac{\ell}{2^{n/2} - \ell} + \delta.$$

Lemma 4. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a MCOE scheme as in Definition 1 (ii). Let $q \leq 2^{n/2-2}$ be the number of total queries an adversary A is allowed to ask and ℓ be an integer representing the total length in blocks of the queries to \mathcal{E} and \mathcal{D} . Then,*

$$\begin{aligned} \mathbf{Adv}_{\Pi}^{\text{INT-CTXT(NI)}}(q, \ell, t) &\leq \frac{(2q + \ell + 2)(2q + \ell + 3)}{2^n - (q + \ell)} + \frac{2(2q + \ell)}{2^n - (q + \ell)} + \frac{q(q + 1)}{2^n - q} \\ &\quad + \frac{q}{2^{n/2} - q} + \mathbf{Adv}_{\tilde{E}, \tilde{E}^{-1}}^{\text{T-IND-CCA}}(2q + \ell, O(t)). \end{aligned}$$

Proof. Our bound is derived by game playing arguments. Consider games G_1 - G_3 of Figure 10 and a fixed adversary A asking at most q queries with a total length of at most ℓ blocks. The functions **Initialize** and **Finalize** are identical for all games in this proof. Lets denote G_0 as the Game INT-CTXT(NI) as defined in Figure 7. Definition 8 states that

$$\mathbf{Adv}_{\Pi}^{\text{INT-CTXT(NI)}}(A) \leq \Pr[A^{G_0} \Rightarrow 1].$$

In G_1 , the encryption and verify placeholders are replaced by their generic MCOE counterparts as of Definition 1. We now discuss the differences between G_1 and G_2 . The set B is initialized with $\{0^n, 1^n\}$ and then collects all new key-input values U which are computed during the encryption or verification process (in lines 209, 215, 227, 240, 246, and 258). Furthermore, the sets $A[U]$ collect the masked values of M^* (cf. lines 220 and 252) for a certain prefix.

In lines 202 and 233, the LLCP_n oracle is inquired. Finally, the variable **bad** is set to **true** if one of the if-conditions in lines 207, 213, 218, 225, 238, 244, 249 or 256 hold. *None* of these modifications affect the values returned to the adversary and therefore

$$\Pr[A^{G_1} \Rightarrow 1] = \Pr[A^{G_2} \Rightarrow 1].$$

For our further discussion we require another game G_4 which is explained in more detail later in this proof³. It follows that

$$\begin{aligned} \Pr[A^{G_2} \Rightarrow 1] &\leq \Pr[A^{G_3} \Rightarrow 1] + |\Pr[A^{G_2} \Rightarrow 1] - \Pr[A^{G_3} \Rightarrow 1]| \\ &\leq \Pr[A^{G_3} \Rightarrow 1] + \Pr[A^{G_3} \text{ sets bad}] \\ &\leq \Pr[A^{G_4} \Rightarrow 1] + |\Pr[A^{G_3} \Rightarrow 1] - \Pr[A^{G_4} \Rightarrow 1]| + \Pr[A^{G_3} \text{ sets bad}]. \end{aligned} \quad (4)$$

We now proceed to upper bound any of the three terms contained in (4) – in right to left order. The success probability of game G_3 does not differ from the success probability of G_2 unless a chaining value U occurs twice. In this case, the adversary must (1) either have ‘found’ a collision for $\tilde{E}_K(X, Y) \oplus Y$, *i.e.*, she stumbles over (X, Y) and (X', Y') such that $\tilde{E}_K(X, Y) \oplus Y = \tilde{E}_K(X', Y') \oplus Y'$ or, (2), must have found a preimage of 0^n or 1^n , which is always the starting point of our chain. Note, the values 0^n and 1^n are initially stored in the set B . In both cases, the variable **bad** would have been set to **true**. From [9] follows as upper bound

$$\frac{(2q + \ell + 2)(2q + \ell + 3)}{2^n - (q + \ell)} + \frac{2(2q + \ell)}{2^n - (q + \ell)}.$$

³ Since the difference is very small, we do not provide an extra figure.

```

1       $K \xleftarrow{\$} \mathcal{K}(); B \leftarrow \{0^n, 1^n\};$ 

100 Encrypt( $H, M$ ) Game  $G_1$ 
101    $L_H \leftarrow |H|/n; L \leftarrow \lceil |M|/n \rceil;$ 
102    $U \leftarrow 0^n;$ 
103   for  $i = 1, \dots, L_H$  do
104      $\tau \leftarrow \tilde{E}_K(U, H_i);$ 
105      $U \leftarrow H_i \oplus \tau;$ 
106   for  $i = 1, \dots, L - 1$  do
107      $C_i \leftarrow \tilde{E}_K(U, M_i);$ 
108      $U \leftarrow C_i \oplus M_i;$ 
109    $M^* \leftarrow M_L || \tau[0, \dots, n - |M_L| - 1];$ 
110    $M^* \leftarrow M^* \oplus \tilde{E}_K(1^n, |M_L|);$ 
111    $C^* \leftarrow \tilde{E}_K(U, M^*);$ 
112    $C_L \leftarrow C^*[0, \dots, |M_L| - 1];$ 
113    $T[0, \dots, n - |M_L| - 1] \leftarrow C^*[|M_L| - 1, \dots, n - 1];$ 
114    $U \leftarrow M^* \oplus C^*;$ 
115    $T[n - |M_L|, \dots, n - 1] \leftarrow \tilde{E}_K(U, \tau)[0, \dots, |M_L|];$ 
116    $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, M, C, T)\};$ 
117   return  $(C_1, \dots, C_L, T);$ 

200 Encrypt( $H, M$ ) Game  $G_2, \boxed{G_3}$ 
201    $L_H \leftarrow |H|/n; L \leftarrow \lceil |M|/n \rceil;$ 
202    $p \leftarrow \text{LLCP}_n(\mathcal{Q}_{|H, M}, (H, M));$ 
203    $U \leftarrow 0^n;$ 
204   for  $i = 1, \dots, L_H$  do
205      $\tau \leftarrow \tilde{E}_K(U, H_i);$ 
206      $U \leftarrow H_i \oplus \tau;$ 
207     if  $(U \in B \text{ and } i > p)$  then
208       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B;$ 
209      $B \leftarrow B \cup U;$ 
210   for  $i = 1, \dots, L - 1$  do
211      $C_i \leftarrow \tilde{E}_K(U, M_i);$ 
212      $U \leftarrow C_i \oplus M_i;$ 
213     if  $(U \in B \text{ and } i + L_H > p)$  then
214       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B;$ 
215      $B \leftarrow B \cup U;$ 
216    $M^* \leftarrow M_L || \tau[0, \dots, n - |M_L| - 1];$ 
217    $M^* \leftarrow M^* \oplus \tilde{E}_K(1^n, |M_L|);$ 
218   if  $(M^* \in A[U] \text{ and } L + L_H - 1 = p)$  then
219     bad  $\leftarrow$  true;  $M^* \xleftarrow{\$} \{0, 1\}^n \setminus A[U];$ 
220    $A[U] \leftarrow A[U] \cup M^*;$ 
221    $C^* \leftarrow \tilde{E}_K(U, M^*);$ 
222    $C_L \leftarrow C^*[0, \dots, |M_L| - 1];$ 
223    $T[0, \dots, n - |M_L| - 1] \leftarrow C^*[|M_L| - 1, \dots, n - 1];$ 
224    $U \leftarrow C^* \oplus M^*;$ 
225   if  $(U \in B \text{ and } L + L_H > p)$  then
226     bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B;$ 
227    $B \leftarrow B \cup U;$ 
228    $T[n - |M_L|, \dots, n - 1] \leftarrow \tilde{E}_K(U, \tau)[0, \dots, |M_L| - 1];$ 
229    $\mathcal{Q} \leftarrow \mathcal{Q} \cup (H, M, C, T);$ 
230   return  $(C_1, \dots, C_L, T);$ 

2   Finalize ()
3   return win;

118 Verify( $H, C, T$ ) Game  $G_1$ 
119    $L_H \leftarrow |H|/n; L \leftarrow \lceil |C|/n \rceil;$ 
120    $U \leftarrow 0^n;$ 
121   for  $i = 1, \dots, L_H$  do
122      $\tau \leftarrow \tilde{E}_K(U, H_i);$ 
123      $U \leftarrow H_i \oplus \tau;$ 
124   for  $i = 1, \dots, L - 1$  do
125      $M_i \leftarrow \tilde{E}_K^{-1}(U, C_i);$ 
126      $U \leftarrow C_i \oplus M_i;$ 
127    $C^* \leftarrow C_L || T[0 \dots n - |C_L|^* - 1];$ 
128    $M^* \leftarrow \tilde{E}_K^{-1}(U, C^*);$ 
129    $U \leftarrow M^* \oplus C^*;$ 
130    $M^* \leftarrow M^* \oplus \tilde{E}_K(1^n, |C_L|);$ 
131    $M_L \leftarrow M^*[0, \dots, |C_L| - 1];$ 
132    $\tau'[0 \dots n - |C_L| - 1] \leftarrow M^*[|C_L|, \dots, n - 1];$ 
133    $T' \leftarrow \tilde{E}_K(U, \tau);$ 
134   if  $\tau'[0 \dots n - l^* - 1] = \tau[0 \dots n - l^* - 1]$ 
135     and  $T'[0 \dots l^* - 1] = T[n - l^* \dots n - 1]$ 
136     and  $(H, C) \notin \mathcal{Q}_{|H, C}$  then
137     win  $\leftarrow$  true;
138    $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, \perp, C, \perp)\};$ 
139   return win;

231 Verify( $H, C, T$ ) Game  $G_2, \boxed{G_3}$ 
232    $L_H \leftarrow |H|/n; L \leftarrow \lceil |C|/n \rceil;$ 
233    $p \leftarrow \text{LLCP}_n(\mathcal{Q}_{|H, M}, (H, M));$ 
234    $U \leftarrow 0^n;$ 
235   for  $i = 1, \dots, L_H$  do
236      $\tau \leftarrow \tilde{E}_K(U, H_i);$ 
237      $U \leftarrow H_i \oplus \tau;$ 
238     if  $(U \in B \text{ and } i > p)$  then
239       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B;$ 
240      $B \leftarrow B \cup U;$ 
241   for  $i = 1, \dots, L - 1$  do
242      $M_i \leftarrow \tilde{E}_K^{-1}(U, C_i);$ 
243      $U \leftarrow C_i \oplus M_i;$ 
244     if  $(U \in B \text{ and } i + L_H > p)$  then
245       bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B;$ 
246      $B \leftarrow B \cup U;$ 
247    $C^* \leftarrow C_L || T[0 \dots n - |C_L|^* - 1];$ 
248    $M^* \leftarrow \tilde{E}_K^{-1}(U, C^*);$ 
249   if  $(M^* \in A[U] \text{ and } L + L_H - 1 = p)$  then
250     bad  $\leftarrow$  true;  $M^* \xleftarrow{\$} \{0, 1\}^n \setminus A[U];$ 
251    $A[U] \leftarrow A[U] \cup M^*;$ 
252    $M^* \leftarrow M^* \oplus \tilde{E}_K(1^n, |M_L|);$ 
253    $M_L \leftarrow M^*[0, \dots, |C_L| - 1];$ 
254    $\tau'[0 \dots n - |C_L| - 1] \leftarrow M^*[|C_L|, \dots, n - 1];$ 
255    $U \leftarrow M^* \oplus C^*;$ 
256   if  $(U \in B \text{ and } L + L_H + 1 > p)$  then
257     bad  $\leftarrow$  true;  $U \xleftarrow{\$} \{0, 1\}^n \setminus B;$ 
258    $B \leftarrow B \cup U;$ 
259    $T' \leftarrow \tilde{E}_K(U, \tau);$ 
260   if  $\tau'[0 \dots n - l^* - 1] = \tau[0 \dots n - l^* - 1]$ 
261     and  $T'[0 \dots l^* - 1] = T[n - l^* \dots n - 1]$ 
262     and  $(H, C) \notin \mathcal{Q}_{|H, C}$  then
263     win  $\leftarrow$  true;
264    $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, \perp, C, \perp)\};$ 
265   return win;

```

Fig. 10. Games G_1 - G_3 for the proof of Lemma 4. Game G_3 contains the code in the box while G_2 does not.

Furthermore, we have to consider the case when a collision occurs between the masked value of M^* and the set $A[U]$. As an adversary can ask q queries, it follows that the probability of the flag **bad** is set to **true** in lines 219 and 250 can be upper bounded by

$$\frac{q(q+1)}{2^n - q}.$$

By adding up both bounds it follows that

$$\Pr[A^{G_3} \text{ sets } \mathbf{bad}] \leq \frac{(2q + \ell + 2)(2q + \ell + 3)}{2^n - (q + \ell)} + \frac{2(2q + \ell)}{2^n - (q + \ell)} + \frac{q(q + 1)}{2^n - q}.$$

Now we describe the new game G_4 , which is equal to G_3 *except* that the block cipher \tilde{E} and its inverse \tilde{E}^{-1} are replaced by the functions **EncryptBlock** and **DecryptBlock**. These are modeled as a set of pseudorandom permutations, where the index is given by the tweak. We assume that they are implemented via lazy sampling. More precisely, the call $\tilde{E}_K(X, Y)$ is replaced by an invocation of **EncryptBlock** $_K(X, Y)$ and the call $\tilde{E}_K^{-1}(X, Y)$ by an invocation of **DecryptBlock** $_K(X, Y)$. In the following, we upper bound the difference between G_3 and G_4 by

$$|\Pr[A^{G_3} \Rightarrow 1] - \Pr[A^{G_4} \Rightarrow 1]| \leq \mathbf{Adv}_{\tilde{E}, \tilde{E}^{-1}}^{\text{T-IND-CCA}}(2q + \ell, O(t)).$$

Finally, we have to upper bound the advantage for the adversary A to win the game G_4 . A can win this game only if the condition in lines 260-262 (resp. 460-462 for game G_4) holds. As usual, we assume *wlog.* that A does not ask a question if the answer is already known. This implies that $(H, C, T) \notin \mathcal{Q}_{H, C, T}$. We formally adjust lines 260-262 (*i.e.*, choose as the tag computation operation either \tilde{E} or \tilde{E}^{-1}) such that we always have enough randomness left for our result. For the sake of simplicity, we denote the two final chaining values by U_L and U_{L+1} . For our analysis, we distinguish between two main cases: the case when $|M_L| = n$, *i.e.*, the size of the last message block is equal to the block size n , and is opposite case.

Case 1: In this case we first consider that $U_L \in B$. This implies that (C_1, \dots, C_L) must be part of a common prefix of a previous query. The adversary can win only if T is new, *i.e.*, not a part of a previously occurred prefix. The upper bound is then given by

$$\Pr \left[\tilde{E}_K^{-1}(U_{L+1}, T) = \tau \right] = 0$$

since \tilde{E}^{-1} is a PRP.

If $U_L \notin B$, we can upper bound the success probability for one query by $1/(2^n - q)$. Hence, for q queries, we can upper bound the success probability by $q/(2^n - q)$.

Case 2: Now, we consider the case $|M_L| \neq n$. It can be upper bounded by Lemma 5. The success probability is at most $q/(2^{n/2} - q)$.

Since both cases are mutually exclusive, we can upper bound the success probability for q queries by

$$\frac{q}{2^{n/2} - q}.$$

Our claim follows by adding up the individual bounds. \square

Lemma 5. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a MCOE scheme as in Definition 2 and q be the number of total queries with $q \leq 2^{n/2-2}$. The probability that an adversary A wins G_4 as defined in the proof of Lemma 4, for any message which is not a multiple of the block size n , can be upper bounded by $q/(2^{n/2} - q)$.*

Proof. In this proof, we denote the two final chaining values as U_L and U_{L+1} . Furthermore, we denote by T^α the $(n - |C_L|)$ -bit string $T[0..n - |C_L| - 1]$ and by T^β the $|C_L|$ -bit string $T[n - |C_L|, \dots, n]$. Additionally, we denote the corresponding strings $\tau^\alpha = \tau[0, \dots, |C_L| - 1]$ and $\tau^\beta = \tau[|C_L|, \dots, n - 1]$, respectively.

Case 1: $U_{L+1} \in B$:

This case implies that $(C_1, \dots, C_L, T^\alpha)$ must be part of a common prefix from a previous query, otherwise, this would imply a collision of the chaining value. But this event is already handled by setting the flag `bad` to `true` in game G_2 (cf. Line 248 of Figure 4). Hence, the adversary can only win if T^β is new, *i.e.*, not a part of a previously occurred prefix. The upper bound is given by

$$\max_Z \left\{ \Pr[\tilde{E}_K^{-1}(U_{L+1}T^\beta || Z) = \tau] \right\} = 0$$

since \tilde{E}^{-1} is a PRP.

Case 2: $U_{L+1} \notin B$:

This case implies that $C_L || T^\alpha$ must be new. The probability that the condition from Line 260 holds – for q queries – can be upper bounded by

$$\Pr_\alpha = \max_{M_L} \left\{ \Pr \left[\tilde{E}_K^{-1}(U_L, C_L || T^\alpha) = (M_L || \tau^\alpha) \oplus \tilde{E}_K(1^n, |M_L|) \right] \right\} \leq \frac{2}{2^{(n-|C_L|)} - 2q}.$$

Hence, the probability for q queries can be upper bounded by $\frac{2q}{2^{(n-|C_L|)} - 2q}$.

From the assumption $U_{L+1} \notin B$ follows that U_{L+1} is new. Since \tilde{E} is a PRP, we can upper bound the probability that the condition from Line 452 holds by

$$\Pr_\beta = \max_Z \left\{ \Pr[\tilde{E}_K(U_{L+1}\tau) = T^\beta || Z] \right\} \leq \frac{1}{2^{|C_L|} - q}.$$

Then, the probability for q queries can be upper bound by $q/(2^{|C_L|} - q)$.

The success probability of this case depends on the length of $|C_L|$. So we can distinguish between the following three subcases:

Subcase 2.1: $|C_L| < n/2$:

In this case, we can upper bound \Pr_α by $\frac{1}{2^{n/2} - q}$ and \Pr_β by 1. Hence, the total success probability for q queries is at most $\frac{q}{2^{n/2} - q}$.

Subcase 2.2: $|C_L| = n/2$:

In this case, we can upper bound \Pr_α by $\frac{2}{2^{n/2} - 2q}$ and \Pr_β by $\frac{1}{2^{n/2} - q}$. Hence, the total success probability for q queries is at most $\frac{2q^2}{2^{n-1} - q^2}$.

Subcase 2.3: $|C_L| > n/2$:

In this case, we can upper bound \Pr_α by 1 and \Pr_β by $\frac{1}{2^{n/2+1} - q}$. Hence, the total success probability for q queries is at most $\frac{q}{2^{n/2+1} - q}$.

Since all three subcases are mutually exclusive, we can upper bound the success probability for $q \leq 2^{n/2-2}$ queries by

$$\max \left\{ \frac{q}{2^{n/2} - q}, \frac{2q^2}{2^{n-1} - q^2}, \frac{q}{2^{n/2+1} - q} \right\} \leq \frac{q}{2^{n/2} - q}.$$

Due to the fact that Case 1 and Case 2 are mutually exclusive, we can upper bound the success probability for q queries by

$$\max \left\{ 0, \frac{q}{2^{n/2} - q} \right\} \leq \frac{q}{2^{n/2} - q}.$$

Our claim follows by adding up the individual bounds. □

Lemma 6. Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a MCOE scheme as in Definition 2. Let q be the number of total queries an adversary A is allowed to ask and ℓ be an integer representing the total length of the queries to \mathcal{E} and \mathcal{D} . Then,

$$\begin{aligned} \mathbf{Adv}_{\Pi}^{\text{CPA}(\text{AOE}, \text{NI})}(q, \ell, t) \leq & 2 \left(\frac{(q + \ell + 2)(q + \ell + 3)}{2^n - (q + \ell)} + \frac{2(2q + \ell)}{2^n - (q + \ell)} + \frac{q(q + 1)}{2^n - q} \right) \\ & + 2\mathbf{Adv}_{\tilde{E}}^{\text{T-IND-CPA}}(2q + \ell, O(t)). \end{aligned}$$

Proof. First, we investigate the differences between the $\text{CPA}(\text{AOE}, \text{NI})$ -game from Figure 6 and G_1 from Figure 11. In G_1 we have replaced \mathcal{E} by its definition of MCOE, and \mathcal{E}^w by an on-line encryption oracle **OPerm** (line 102) that just models a 'perfect' OPRP, *i.e.*, for two plaintexts with an equal prefix it returns two ciphertexts that also share a prefix of the same length. Again we assume this oracle to be implemented by lazy sampling. Then, set B collects all chaining values (lines 113 and 119) in order to intercept the occurrence of two equal chaining values which lead to two identical tweaks for the encryption of a block. The sets $A[U]$ collect all values of masked M^* for specific chaining values U (line 124).

In Line 105, the oracle LLCP_n is invoked returning the length of the longest common prefix of (H, M) and $\mathcal{Q}_{|H, M}$. Finally, the variable **bad** is set to **true** if (one of) the conditions of lines 111/211, 117/217, or 122/222 holds. These changes do not affect the success probability of an adversary because the oracle output remains unchanged. More precisely, the distribution of the output does not change. This means that

$$\mathbf{Adv}_{\Pi}^{\text{CPA}(\text{AOE}, \text{NI})}(A) = 2 \cdot |\Pr[A^{G_1} \Rightarrow 1] - 0.5|,$$

and therefore, by common game-playing arguments – using a new game G_3 described shortly –

$$\begin{aligned} \Pr[A^{G_1} \Rightarrow 1] & \leq \Pr[A^{G_2} \Rightarrow 1] + |\Pr[A^{G_1} \Rightarrow 1] - \Pr[A^{G_2} \Rightarrow 1]| \\ & \leq \Pr[A^{G_2} \Rightarrow 1] + \Pr[A^{G_2} \text{ sets bad}] \\ & \leq \Pr[A^{G_3} \Rightarrow 1] + |\Pr[A^{G_2} \Rightarrow 1] - \Pr[A^{G_3} \Rightarrow 1]| + \Pr[A^{G_2} \text{ sets bad}]. \end{aligned}$$

The success probability of game G_2 does not differ from the success probability of G_1 unless (1) a chaining value U occurs twice or (2) a collision between two masked values M^* – sharing the same chaining value U – occurs.

In the first case, the adversary must either have found a collision for $\tilde{E}_K(X, Y) \oplus Y$, *i.e.*, she has found (X, Y) and (X', Y') such that $\tilde{E}_K(X, Y) \oplus Y = \tilde{E}_K(X', Y') \oplus Y'$, or a preimage of 0^n or 1^n . In these cases, the variable **bad** would have been set to **true**, and it follows again by [9] that

$$\frac{(q + \ell + 2)(q + \ell + 3)}{2^n - (q + \ell)} + \frac{2(2q + \ell)}{2^n - (q + \ell)}$$

In the second case, an adversary can ask q queries, and it follows that the probability for the flag **bad** is set to **true** in Line 123 can be upper bounded by

$$\frac{q(q + 1)}{2^n - q}.$$

By adding up both bounds follows

$$\Pr[A^{G_2} \text{ sets bad}] \leq \frac{(q + \ell + 2)(q + \ell + 3)}{2^n - (q + \ell)} + \frac{2(2q + \ell)}{2^n - (q + \ell)} + \frac{q(q + 1)}{2^n - q}.$$

The aforementioned new game G_3 is equal to the game G_2 *except* that the block cipher \tilde{E} and its inverse \tilde{E}^{-1} are replaced by randomly chosen functions **EncryptBlock** and **DecryptBlock**, which are modeled as pseudo-random permutations. We assume that they are

<pre> 1 Initialize () 2 $b \xleftarrow{\\$} \{0,1\}; K \xleftarrow{\\$} \mathcal{K}(); B \leftarrow \{0^n, 1^n\};$ </pre>	<pre> 3 Finalize (d) 4 return (b=d); </pre>
<pre> 100 Encrypt(H, M) Game $G_1, \boxed{G_2}$ 101 if ($b=0$) then 102 $C \leftarrow \mathbf{OPerm}(H, M);$ 103 else 104 $L_H \leftarrow H /n; L \leftarrow M /n;$ 105 $p \leftarrow \text{LLCP}_n(\mathcal{Q}, (H, M));$ 106 $\mathcal{Q} \leftarrow \mathcal{Q} \cup (H, M);$ 107 $U \leftarrow 0^n;$ 108 for $i=1, \dots, L_H$ do 109 $\tau \leftarrow \tilde{E}_K(U, H_i);$ 110 $U \leftarrow H_i \oplus \tau;$ 111 if ($U \in B$ and $i > p$) then 112 $\text{bad} \leftarrow \text{true}; \boxed{U \xleftarrow{\\$} \{0,1\}^n \setminus B};$ 113 $B \leftarrow B \cup U;$ </pre>	<pre> 114 for $i=1, \dots, L-1$ do 115 $C_i \leftarrow \tilde{E}_K(U, M_i);$ 116 $U \leftarrow C_i \oplus M_i;$ 117 if ($U \in B$ and $i + L_H > p$) then 118 $\text{bad} \leftarrow \text{true}; \boxed{U \xleftarrow{\\$} \{0,1\}^n \setminus B};$ 119 $B \leftarrow B \cup U;$ 120 $M^* \leftarrow M_L \tau[0, \dots, n - M_L - 1];$ 121 $M^* \leftarrow M^* \oplus \tilde{E}_K(1^n, M_L);$ 122 if ($M^* \in A[U]$ and $L + L_H - 1 = p$) then 123 $\text{bad} \leftarrow \text{true}; \boxed{M^* \xleftarrow{\\$} \{0,1\}^n \setminus A[U]};$ 124 $A \leftarrow A[U] \cup M^*;$ 125 $C^* \leftarrow \tilde{E}_K(U, M^*);$ 126 $C_L \leftarrow C^*[0, \dots, M_L - 1];$ 127 return (C_1, \dots, C_L); </pre>

Fig. 11. Games G_1 and G_2 for the proof of Lemma 3. Game G_2 contains the code in the box while G_1 does not.

implemented via lazy sampling. More precisely, the call $\tilde{E}_K(X)$ is replaced by an invocation of **EncryptBlock** $_K(X)$ and the call $\tilde{E}_K^{-1}(X)$ is replaced by an invocation of **DecryptBlock** $_K(X)$. We now upper bound the difference between G_2 and G_3 . So, by definition of G_4 , we have

$$|\Pr[A^{G_2} \Rightarrow 1] - \Pr[A^{G_3} \Rightarrow 1]| \leq \mathbf{Adv}_{\tilde{E}, \tilde{E}^{-1}}^{\text{T-IND-CPA}}(2q + \ell, O(t)).$$

Finally, we have to upper bound the advantage for an adversary A to win the game G_3 . Since U cannot collide and it is not possible to compute a preimage for any query, the algorithm for $b = 0$ is an OPRP, and therefore the success probability to win G_3 for any adversary is 0.5, *i.e.*, she has no advantage in winning this game.

Our claim follows by adding up the individual bounds. \square

6 The On-Line Authenticated Encryption Scheme McOE-X

This section shows the security of McOE-X for any block cipher. The here-presented upper bounds are based on the generic proof of the McOE family as showed in Section 5. Note, that we have generalized the xor-operation between the key and the tweak by a function $\varphi : D_k \times D_n \rightarrow D_n$. For any fixed key K , $\varphi(K, \cdot)$ and the xor-operation are injective. Therefore, we can replace the xor-operation by the function φ .

Theorem 4. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a McOE-X scheme as in Definition 1 and 2, where the tweakable block cipher \tilde{E} is given by*

$$\tilde{E}_K(U, M) := E_{\varphi(K, U)}(M),$$

function. Furthermore, the amount of queries an adversary is allowed to ask is at most $2^{n/2-2}$. Then, the upper bounds for the variants with and without tag-splitting are as follows.

(i) *Security without Tag-Splitting.*

$$\mathbf{Adv}_{\Pi}^{\text{CCA3(OAE,NI)}}(q, \ell, t) \leq \frac{3(q + \ell)(q + \ell + 1) + 4q + 3\ell}{2^n - (q + \ell)} + 3\mathbf{Adv}_E^{\text{RK-CCA-PRP}}(2q + \ell, O(t)).$$

(ii) *Security with Tag-Splitting.*

$$\begin{aligned} \mathbf{Adv}_{\Pi}^{\text{CCA3(OAE,NI)}}(q, \ell, t) \leq & \frac{4(q + \ell + 2)(q + \ell + 3) + 6(2q + \ell)}{2^n - (q + \ell)} + \frac{3q(q + 1)}{2^n - q} + \frac{q}{2^{n/2} - q} \\ & + 3\mathbf{Adv}_E^{\text{RK-CCA-PRP}}(2q + \ell, O(t)). \end{aligned}$$

Proof. The proofs of (i) and (ii) follow from Theorem 2 and Theorem 3. Since φ is an injective function, a collision for the chaining values U implies a collision for the values of $\varphi(K, U)$. Furthermore, it is easy to see that the advantage for the tweakable block cipher can be upper bounded by the RK-CCA-PRP-advantage of an adversary A , even though she has only limited control over the tweak, *i.e.*, the chaining value U . \square

Key Recovery. In [?], Mendel *et al.* introduced a key-recovery attack on MCOE-X with a birthday-bound complexity. The adversary is allowed to query the MCOE-X encryption oracle \mathcal{E} and access the block cipher E itself. In general, this is a reasonable assumption since we can assume E to be a common block cipher like AES. The attacks works as follows:

1. Choose an arbitrary value a .
2. Compute $b_i = E(k_i, a)$ for $i = 1, \dots, q$ and save the pair (b_i, k_i) in a list L .
3. Choose an arbitrary constant value m_1 , set $m_2 = a$, and choose an arbitrary nonce value V_i . Thus, set $M = m_1 \parallel m_2$ and ask for $C = \mathcal{E}(V_i, M)$ with $C = c_1 \parallel c_2$.
4. Check if $c_2 \in L|_b$.
 - 4.1 If so, compute the secret key by $K = k_i \oplus m_1 \oplus c_1$.
 - 4.2 Otherwise, go back to Step 3.

Let us denote q' as the number of iteration for the loop described in steps 3 and 4. The total number of block-cipher invocations ℓ is restricted to $\ell \leq q + 4q'$ since for one query to \mathcal{E} four block-cipher calls are necessary to compute the pair (C, T) . If we choose $\ell \approx 2^{n/2}$, there exists an adversary with a success probability of at most $1/2$, which is able to recover K using the presented attack. Note that this observation is not only in accordance to our security claim, but also provides an inferior security bound. Nevertheless, this attacks shows, that it is absolutely essential to change the key from time to time. We claim that key recovery attacks with a birthday bound complexity are not applicable to MCOE-G. Furthermore, one can avoid these attacks by increasing the key size to, *e.g.*, $2n$.

7 The On-Line Authenticated-Encryption Scheme McOE-G

This section shows the security of MCOE-G for any given block cipher when using an ϵ -AXU-secure hash function. The here-presented upper bounds are based on (1) the generic proof of the MCOE family as showed in Section 5 and (2) the paper of Liskov *et al.* (see Theorem 2 of [31]).

Liskov *et al.* showed that the T-IND-CCA advantage of an adversary A is at most

$$\mathbf{Adv}_{\tilde{E}, \tilde{E}^{-1}}^{\text{T-IND-CCA}}(q, t) \leq \mathbf{Adv}_{E, E^{-1}}^{\text{CCA-PRP}}(q, O(t)) + 3\epsilon q^2 \quad (5)$$

We use this result for the following security proof.

Theorem 5. *Let $\Pi = (\mathcal{K}, \mathcal{E}, \mathcal{D})$ be a MCOE-G scheme as in Definition 1 and 2, where the tweakable block cipher \tilde{E} is given by*

$$\tilde{E}_K(U, M) := E_{K_1}(M \oplus H_{K_2}(U)) \oplus H_{K_2}(U).$$

with $E \in \text{Block}(n, n)$ and H is a family of ϵ -AXU hash functions. Furthermore, the amount of queries an adversary is allowed to ask is at most $2^{n/2-2}$. Then, the upper bounds for the variants with and without tag-splitting method are as follows.

(i) *Security without Tag-Splitting.*

$$\begin{aligned} \mathbf{Adv}_{\Pi}^{\text{CCA3}(\text{OAE,NI})}(q, \ell, t) &\leq \frac{3(q + \ell)(q + \ell + 1) + 4q + 3\ell}{2^n - (q + \ell)} \\ &\quad + 3 \left(\mathbf{Adv}_{E, E^{-1}}^{\text{CCA-PRP}}(2q + \ell, O(t)) + 3\epsilon(2q + \ell)^2 \right) \end{aligned}$$

(ii) *Security with Tag-Splitting.*

$$\begin{aligned} \mathbf{Adv}_{\Pi}^{\text{CCA3}(\text{OAE,NI})}(q, \ell, t) &\leq \frac{4(q + \ell + 2)(q + \ell + 3) + 6(2q + \ell)}{2^n - (q + \ell)} + \frac{3q(q + 1)}{2^n - q} + \frac{q}{2^{n/2} - q} \\ &\quad + 3 \left(\mathbf{Adv}_{E, E^{-1}}^{\text{CCA-PRP}}(2q + \ell, O(t)) + 3\epsilon(2q + \ell)^2 \right) \end{aligned}$$

Proof. The proofs of (i) and (ii) follow from Theorem 2, Theorem 3 and Equation 5. \square

Remark. MCOE-G is not secure if an adversary has oracle access to the internal building blocks, *i.e.*, the block cipher E and the ϵ -AXU-secure hash function H . This is shown by Black *et al.* in [8]. Hence, it is crucial that the adversary may query only the tweakable block cipher \tilde{E} and not one of its parts.

8 Discussion

New Challenges for Research. At this point of time, cryptographic research has developed an impressive number of good schemes for encryption, authentication, and authenticated encryption. Many of these schemes have been proven secure under standard assumptions on the underlying primitives. In practice, however, such schemes are often used in a way that undermines security. Trying to design cryptographic schemes as “misuse-resistant” as possible still remains as a challenge for cryptographers.

Furthermore, our research seems to pose new challenges for the design of symmetric primitives. Ideally, we would like to implement MCOE using a tweakable n -bit block cipher with n -bit tweaks, supports fast random tweak changes. Due to the current lack of such a primitive, we designed MCOE-X, which requires an ordinary n -bit block cipher being secure against XOR-related key attacks, and supporting fast random key changes. Beyond MCOE, cryptosystem designers could benefit from new tweak-agile tweakable block ciphers and new key-agile ordinary block ciphers.

It is mentionable that MCOE-X, when using Threefish-512 in software, performs considerably better as when using software or even hardware-AES-128. Note, Threefish-512 actually is a tweakable block cipher, but the 128-bit tweak is too short for MCOE. As an alternative, we developed further variants of MCOE using double encryption and Galois Field arithmetic. Both variants also do not expose the underlying block cipher to related-key attacks.

Conclusion. Originally, this research has been inspired by the search for a default authenticated encryption mode of operation for a general-purpose cryptographic library. It should offer, by default, a huge failure-tolerance for practical software developers and still allow being used in an on-line manner.

Since the well-known schemes, such as OCB and SIV, did not fit our requirements, we searched for novel ways to achieve the security and functionality we were looking for. Apart from MCOE, generic composition (Encrypt-then-Mac) of a secure on-line cipher for encryption and a secure deterministic MAC for authentication, using two independent keys, might be another solution. As it turned out, using MCOE, one can save the additional key and the time to generate the MAC by using a slightly tweaked on-line cipher for both encryption and authentication.

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Appendices

A Misuse-Attacks: The Weak Point of Current Authenticated Encryption (AE) Schemes.

A.1 Attacking Schemes without Claimed Resistance Against Nonce-Reuse

Cipher-block-chaining (CBC) is an unauthenticated encryption mode which is sometimes used as the encryption component of an AE scheme. But one can easily distinguish CBC encryption from a good on-line cipher, if the nonce (or the IV) is constant. The attack from [1] only needs three chosen plaintexts. Counter mode, which has been very popular among the designers of AE schemes, fails terribly in nonce-reuse settings, since it generates exactly the same keystream twice when a nonce is reused. It was to be expected that a scheme using counter mode or CBC inherits the nonce-reuse issue from that mode. But, as it turned out, common AE schemes also fail at the authenticity frontier (see Table 2 in Section 1 for an overview). This is an unpleasant surprise, since the cryptographic community has known well deterministic MACs for a long time – so why is the authenticity provided by most authenticated encryption schemes so much more fragile than the authenticity provided by well-known MACs?

The following two attack patterns will be used in most of our attacks.

Repeated Keystream. Many AE schemes generate a keystream $S = F_K(V)$ of length $|M|$, depending on the secret key K and the nonce V . They encrypt a message M by computing the corresponding ciphertext $C = S \oplus M$, typically by applying a block cipher in counter mode. If the same nonce is used more than once, the following attack straightforwardly breaks the privacy:

1. Encrypt a plaintext M under the nonce V to a ciphertext C with tag T .
2. Encrypt a plaintext $M' \neq M$ under the same V to a ciphertext C' and a tag T' .
3. It turns out that $C' = C \oplus M \oplus M'$ holds.

Linear Tag. Many AE schemes, which generate a keystream $S = F_K(V)$ as above, apply the Encrypt-then-Mac (EtM) paradigm and allow to rewrite the authentication tag T as

$$T = f(V) \oplus g(C),$$

where V is the nonce, C is the ciphertext, and f and g are some key-dependent functions. This enables the adversary to mount the linear attack introduced by the following four steps.

1. Encrypt the plaintext M under the nonce V to (C, T) with $T = f(V) \oplus g(C)$.
2. Encrypt the plaintext $M' \neq M$ with $|M'| = |M|$ under the nonce $V' \neq V$ to (C', T') with the tag $T' = f(V') \oplus g(C')$.
3. Set $M'' := M' \oplus C' \oplus C$. Encrypt M'' under the nonce V' to (C'', T'') . Observe $C'' = C$, thus $T'' = f(V') \oplus g(C)$.
4. Set $T^* = T \oplus T' \oplus T'' = f(V) \oplus g(C')$, The adversary accepts (C', T^*) under V .

Two-Pass AE(AD) Modes: CWC [29], GCM [34], CCM [14], EAX [7], CHM [22]. All the common two-pass AE(AD) modes, CHM, CWC, GCM, CCM and EAX, use the counter mode as the underlying encryption operation and are thus vulnerable to the repeated keystream attack pattern. Four of them, CHM, CWC, GCM, and EAX, are designed according to the EtM paradigm, and are thus vulnerable to the linear tag attack pattern. The designers of CCM followed Mac-then-Encrypt (MtE), which seems to defend against the linear tag pattern. Forgery attacks against CCM have been presented in [15], though.

Mixed AE(AD) Modes: RPC [11] and CCFB [33]. RPC combines counter mode and electronic codebook mode. Given an n -bit block cipher E under a key K and a c -bit counter cnt , RPC takes an $(n-c)$ -bit plaintext block M_i and computes the ciphertext block $C_i := E_K(M_i || (\text{cnt} + i) \bmod 2^c)$. Authentication is performed locally for each ciphertext block: During decryption, RPC computes $(M_i || X_i) = E_K^{-1}(C_i)$ and accepts M_i as authentic if and only if $X_i = (\text{cnt} + i) \bmod 2^c$. The nonce defines cnt .

Under nonce-reuse, the same sequence $(\text{cnt} + i) \bmod 2^c$ of counter values is used for different messages. This makes it easy to attack the privacy – essentially, when encrypting messages of m $(n-c)$ -bit blocks, RPC degrades into m independent electronic codebooks. Also, given two authentic ciphertexts, (C_1^0, \dots, C_L^0) and (C_1^1, \dots, C_L^1) , any ciphertext $(C_1^{\sigma(1)}, \dots, C_L^{\sigma(L)})$ with $\sigma(i) \in \{0, 1\}$ is valid, since authenticity is verified locally for each $C_i^{\sigma(i)}$.

Similarly to RPC, CCFB is a combination of Counter and CFB mode. Given an $(n-c)$ -bit nonce and $(n-c)$ -bit plaintext blocks M_1, \dots, M_m CCFB generates $(n-c)$ -bit temporary values D_i , c -bit temporary tags T_i and $(n-c)$ -bit ciphertext blocks C_i as follows:

1. $C_0 := N$;
2. for $i \in \{2, \dots, m\}$ do $(D_i || T_i) := E_K(C_{i-1} || \langle i \rangle)$; $C_i := M_i \oplus D_i$
3. $(*, T_{m+1}) := E_K(C_m || \langle m+1 \rangle)$;

Unlike RPC, CCFB only uses the local tags T_i temporarily; the final authentication tag is $T = T_1 \oplus T_2 \oplus \dots \oplus T_{m+1}$.

Note that the first ciphertext block C_1 is essentially the encryption of M_1 in counter mode. Thus, a variant of the repeated keystream pattern is applicable to CCFB. And the following variant of the linear tag pattern applies to CCFB (for simplicity, we assume single-block messages only):

1. Encrypt the plaintext M_1 under V to (C_1, T) .
2. Encrypt the plaintext $M'_1 \neq M_1$ under $V' \neq V$ to (C'_1, T') .
3. Set $M''_1 := M'_1 \oplus C'_1 \oplus C_1$. Encrypt M''_1 under V' to (C''_1, T'') . Observe $C''_1 = C_1$.
4. The adversary accepts $(C'_1, T \oplus T' \oplus T'')$ under V .

One-Pass AE(AD) Modes: IAPM [26], OCB1[40], OCB2[38], OCB3[30], TAE [31]. Given a nonce V and a secret key K , IAPM [26] encrypts a plaintext (M_1, \dots, M_m) to a ciphertext (C_1, \dots, C_m) and an authentication tag T as follows.

Initial step: Generate $m+2$ values s_0, s_1, \dots, s_{m+1} depending on V and K , but not on the plaintext (M_1, \dots, M_m) .

Encryption: For $i \in \{1, \dots, m\}$: $C_i := E_K(M_i \oplus s_i) \oplus s_i$.

Authentication tag: $T := E_K(s_{m+1} \oplus \sum_{1 \leq i \leq m} M_i) \oplus s_0$.

When encrypting messages of m n -bit blocks, IAPM behaves like a set of m independent electronic codebook and, like RPC, is vulnerable to distinguishing attacks based on this. Similarly to RPC, IAPM behaves like a set of m independent electronic codebooks and is vulnerable to the same distinguishing attack. A forgery can exploit the fact that two different same-length messages (M_1, \dots, M_m) and (M'_1, \dots, M'_m) , encrypted under the same nonce, have the same authentication tag $T = E_K(s_{m+1} \oplus \sum_{1 \leq i \leq m} M_i) \oplus s_0 = E_K(s_{m+1} \oplus \sum_{1 \leq i \leq m} M'_i) \oplus s_0$ if $\sum_{1 \leq i \leq m} M_i = \sum_{1 \leq i \leq m} M'_i$.

As much as our attacks are concerned, OCB1–3 and TAE are quite similar to IAPM, and the attacks are the same.

More One-Pass Modes: IACBC [26] and XCBC [17]. Given a nonce V and a secret key K , IACBC [26] encrypts (M_1, \dots, M_m) to (C_1, \dots, C_m) and an authentication tag T as follows.

Initial step: Generate $m + 1$ values s_0, s_1, \dots, s_m depending on V and K , but not on the plaintext (M_1, \dots, M_m) .

Encryption: $x_0 := V$; For $i \in \{1, \dots, m\}$: $x_i := E_K(M_i \oplus x_{i-1})$, $C_i := x_i \oplus s_i$.

Authentication tag: $T := E_K(x_m \oplus \sum_{1 \leq i \leq m} M_i) \oplus s_0$.

Note that under nonce-reuse the authentication tag leaks information about the message. Namely, if we encrypt two messages (M_1, \dots, M_m) and (M'_1, \dots, M'_m) of the same length m under the same nonce, the two authentication tags are the same if and only if $\sum_{1 \leq i \leq m} M_i = \sum_{1 \leq i \leq m} M'_i$.

The following nonce-reuse attack distinguishes IACBC encryption from an online permutation and also provides an existential forgery. For simplicity, we only consider 1-block messages $V \neq W$, which we also use as nonces:

1. Encrypt W under V to (C_1, T) .
2. Encrypt V under W to (C'_1, T') .
3. Encrypt V under V to (C''_1, T'') .
4. Set $C'''_1 := C_1 \oplus C'_1 \oplus C''_1$ and $T''' := T \oplus T' \oplus T''$.
 (C'''_1, T) is a valid encryption of W under W .

Given a nonce V and secret keys K and K' , XCBC encrypts a plaintext (M_1, \dots, M_m) to a ciphertext (C_1, \dots, C_m) and an authentication tag T as follows.

Initial step: Generate $m+1$ values s_1, \dots, s_{m+1} depending on V and K , but not on the plaintext (M_1, \dots, M_m) .

Encryption:

1. $C_0 := E_K(V)$; $x_0 := E_{K'}(V)$;
2. Generate an additional message word $M_{m+1} := x_0 \oplus M_1 \oplus \dots \oplus M_m$ for authentication.
3. For $i \in \{1, \dots, m+1\}$: $x_i := E_K(M_i \oplus x_{i-1})$, $C_i := (x_i + s_i) \bmod 2^n$.

The best attack we have found for XCBC is not quite as baneful as the attacks on the other schemes, as the attack workload is at $O(2^{n/4})$, and the attack only provides a distinguisher, not a forger. For this reuse-nonce chosen-plaintext attack, we ignore the authentication tag:

1. Generate $2^{n/4}$ encryptions of messages M_1^i under a nonce V to C_1^i .
Statistically, expect one pair $i \neq j$ such that the least significant $n/2$ bits of C_1^i are identical to the least significant $n/2$ bits of C_1^j .
2. Generate $2^{n/4}$ encryptions of messages (M_1^i, M_2^k) and (M_1^j, M_2^ℓ) under V to (C_1^i, C_2^k) and (C_1^j, C_2^ℓ) , where the least significant $n/2$ bits of M_2^k and M_2^ℓ are the same.
Statistically, expect one pair $k \neq \ell$ such that $C_2^k = C_2^\ell$ holds.
3. Choose an arbitrary M_3 . Encrypt (M_1^i, M_2^k, M_3) and (M_1^j, M_2^ℓ, M_3) under V to $(C_1^i, C_2^k, C_3^{i,k})$ and $(C_1^j, C_2^\ell, C_3^{j,\ell})$.
Observe $C_3^{i,k} = C_3^{j,\ell}$.

A.2 Dedicated Online Ciphers and Authenticated Encryption

The current paper refers to online ciphers to define the privacy of AOE against general adversaries. Online ciphers have first been studied in [1]. The eprint version of the same paper describes three AE modes for online cipher, using a random nonce V :

1. Prepend V to the plaintext, append redundancy (e.g., a fixed number of 0-bits).
2. Prepend V and then the length of the plaintext, append redundancy.

3. Prepend a random value V as the IV, and append the same V .

The first is secure if and only if the message length is fixed. The second is an obvious repair of the first, but the exact plaintext length must be known at the beginning of the encryption process. This, though using an underlying online cipher, the scheme itself isn't online. The third is vulnerable to a chosen plaintext attack using a message $M||V$, if the adversary can guess V .

A.3 Offline Schemes, Defeating Nonce-Reuse (SIV [41], HBS [25], BTM [24])

Given a nonce N , a message M and associated Data H , these schemes perform two steps:

1. Generate the authentication tag T from H , M , and N .
2. Encrypt M in counter mode, using T as the nonce.

This is inherently offline, because one must finish step 1 before one can start step 2. All of SIV, HBS, and BTM perform counter mode encryption, but employ different MAC schemes to generate the tag T .

This usage of the counter mode is vulnerable in an *on-line decryption misuse* case, where, during decryption, a would-be plaintext is compromised before the tag has been verified. A chosen-ciphertext adversary can exploit that to determine an unknown keystream and then to decrypt an unknown message.

Another misuse case may apply when nonce-reuse is possible and the sender reads the message twice, once for each of the two steps – if there is any chance that *the message has been modified* between the two read operations.

Note that both misuse cases become quite harmless if one replaces the counter mode encryption by the application of an on-line permutation.

B Proof of Theorem 1

Consider games G_0, G_1, G_2 of Figure 12. For a fixed $\text{CCA3}(\omega, \nu)$ adversary A on the scheme Π it holds that

$$\begin{aligned} \Pr[A_{\Pi}^{\text{CCA3}(\omega, \nu)} \Rightarrow 1] &= \Pr[A^{G_0} \Rightarrow 1] \\ &= \Pr[A^{G_1} \Rightarrow 1] + (\Pr[A^{G_0} \Rightarrow 1] - \Pr[A^{G_1} \Rightarrow \text{true}]) \\ &\leq \Pr[A^{G_1} \Rightarrow 1] + \Pr[A^{G_1} \text{ sets bad}]. \end{aligned}$$

Since the **Decrypt** oracles of G_1 and G_2 always return \perp ,

$$\Pr[A^{G_1} \Rightarrow 1] = \Pr[A^{G_2} \Rightarrow 1].$$

Now, we design two adversaries A_c and A_p so that

$$\begin{aligned} \Pr[A^{G_1} \text{ sets bad}] &\leq \Pr[A_c \stackrel{\text{INT-CTXT}(\omega, \nu)}{\Pi} \Rightarrow 1] \text{ and} \\ \Pr[A^{G_2} \Rightarrow 1] &\leq \Pr[A_p \stackrel{\text{CPA}(\omega, \nu)}{\Pi} \Rightarrow 1]. \end{aligned}$$

A_p : Adversary A_p simply runs A answering A 's **Encrypt** queries using its own **Encrypt** oracle, and answers **Decrypt** queries with \perp . A_p outputs whatever A outputs.

A_c : Adversary A_c runs A answering A 's **Encrypt** queries using its own **Encrypt** oracle. It submits A 's **Decrypt** queries to its **Verify** oracle (*cf.* Figure 7) and, regardless of the response, returns \perp . Note that the **Verify** oracle sets **win** to **true** if and only if a fresh **Decrypt** query is valid. Just such a query would set the variable **bad** to **true**. \square

<pre> 1 Initialize(ω, ν) 2 $b \xleftarrow{\\$} \{0, 1\}$; 3 if ($b=1$) then 4 $K \leftarrow \mathcal{K}()$; 5 Finalize($d$) 6 return ($d=b$); </pre>	<pre> 7 Encrypt(H, M) 8 if ($\nu = \text{NR}$ and $V \in B$) then 9 return \perp; 10 else 11 $B \leftarrow B \cup \{V\}$; 12 if ($b=1$) then 13 $C \leftarrow \mathcal{E}_K(H, M)$; 14 else 15 $C \leftarrow \mathcal{S}^\omega(H, M)$; 16 $\mathcal{Q} \leftarrow \mathcal{Q} \cup \{(H, C)\}$; 17 return C; </pre>	<pre> 100 Decrypt(H, C) Game $G_0, \boxed{G_1}$ 101 $M \leftarrow \perp$; 102 if ($(H, C) \notin \mathcal{Q}$ and $b=1$) then 103 $M \leftarrow \mathcal{D}_K(H, C)$; 104 if ($M \neq \perp$) then 105 bad \leftarrow true; $M \leftarrow \perp$; 106 return M; 200 Decrypt(H, C) Game G_2 201 return \perp; </pre>
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Fig. 12. Games G_0, G_1 and G_2 for the proof of Theorem 1. Game G_1 contains the code in the box while G_0 does not. H_0 denotes the first block of the header representing the nonce/initial value.